

# New Directions of Modern Cryptography



**Zhenfu Cao**



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# Preface

The study of cryptography is motivated by security requirements in the real world and also driven forward by security requirements. Regardless of its real applications, cryptography can be viewed as a branch of mathematics. All the new directions of modern cryptography introduced in this book, including proxy reencryption, attribute-based cryptography, batch cryptography, and noncommutative cryptography, have arisen from the requirements. In this book, we focus on the fundamental definitions, precise assumptions, and rigorous security proofs of cryptographic primitives and related protocols, as well as how they developed from the security requirements and how they are applied.

As we know, modern cryptography has evolved dramatically since the 1970s. Nowadays, the field of cryptography encompasses much more than secure communication. It covers, for example, authentication, digital signature, key establishment and exchange, zero-knowledge, secure multiparty computation, electronic auction and election, digital cash, access control, etc. Modern cryptography is concerned with security problems that arise in a variety of distributed environments where attacks may come from either internal or external forces. Instead of giving a rigid and perfect definition of modern cryptography, we here say that, from the viewpoint of applications, modern cryptography is the science and technology that focuses on defending digital information storage, transportation, and distributed computation via modern communication networks.

These networks consist of but are not limited to wired or wireless telecommunication networks, satellite communication networks, broadcast and TV networks, computer networks (including organization-wide intranet and the Internet), and all newly emerging networks, such as the internet of things, cloud computing, social networks, and named data networks. Another important difference between classic cryptography and modern cryptography is related to who is using it. Historically, the major users of cryptography have been military and intelligence organizations. Today, however, cryptography is required everywhere in our lives. Security mechanisms that rely on cryptography become an essential ingredient of information systems. For example, cryptographic methods are used to enforce access control to all web sites and to prevent adversaries from extracting business secrets from stolen laptops. In view of increasing demands on the network security, this book presents some application paradigms and general principles regarding new directions of modern cryptography.

In short, modern cryptography has gone from an art that deals with secret communication for the military and governments to the science and technology that help ordinary people to set up secure systems. This also means that cryptography becomes a more and more central topic within computer science.

In fact, with the rise of new network architectures and services, the security requirements have changed significantly, that is, from single-user communication (each side is of a single user) to multiuser communication (at least one side is of multiple users). The public-key cryptosystem proposed by Diffie and Hellman in 1976 is not sufficient to satisfy the security requirements in multiuser settings. Under the new environments of “one sender vs. multiple receivers” (one-to-many), “multiple senders vs. one receiver” (many-to-one), “multiple senders vs. multiple receivers” (many-to-many), it is a tendency to design and analyze the multiuser-oriented cryptographic algorithms. They aim to solve ciphertext access control problems, trust problems, efficiency problems in the multimessage cryptology, and challenging problems of quantum and biological com-

puting etc.

In the last 10 years, I, as a founder and a director of Trusted Digital Technology Laboratory (TDT Lab) of Shanghai Jiao Tong University, have witnessed the progressively increased demands on cryptographic techniques. I am proud of having been engaged in cryptography research for over 30 years. The TDT Lab is one of the earliest groups focusing publicly on “Trusted-X Technology” in the world. The TDT Lab focuses on the research of cryptology and trusted digital technology. In cryptology, our research interests mainly include authorized cryptography (proxy cryptography, proxy re-cryptography), attribute-based cryptography (identity-based cryptography, spatial cryptography, functional cryptography), post-quantum cryptography (non-commutative cryptography, lattice-based cryptography), collaboration cryptography (aggregated cryptography, batch cryptography), biologic cryptography (DNA cryptography, biometric feature based cryptography), commitment and zero-knowledge proof, as well as cryptanalysis. In trusted digital technology, we mainly study trusted computing, trusted networks, secure storage/access, secure e-commerce/e-government, key management, and so on. The TDT Lab tries to pursue original innovation in fundamental research, acquire intellectual properties in key technology, and promote industrial development inspired by our academic results. During these years, I have instructed 2 post-doctorals, supervised 19 Ph.D.s and 50 masters, and obtained a number of interesting results (along with my students and colleagues). This book can be viewed as a part of the collection of these results.

## **Audience**

The goal of this book is to be of interest to cryptographers and practitioners of network security. In particular, it is aimed at the following readers: For students who complete first degree courses in computer/information science or applied mathemat-

ics and plan to pursue a degree or career in network security, this book may serve as an advanced course in applied cryptography. Fresh Ph.D. candidates beginning their research in cryptography or information/network security would appreciate the new directions introduced in this book. For security researchers and engineers who are interested in the cloud computing security, e-health security, vehicular ad-hoc network security, RFID security, delay tolerant network security, network coding security, and other wired/wireless network security, and who are responsible for designing and developing secure network systems, this book may help them have a solid understanding of the security principles and correctly deploy the applications. This book is also designed to serve as a reference for graduate courses in cryptography, computer sciences, and mathematics, or as a general introduction suitable for people who want to learn cryptography and security themselves.

# Acknowledgments

I would like to acknowledge all those who helped in the development of this book.

First, I am deeply grateful to Professor Xiaolei Dong for devoting her time to improving this book with invaluable comments, criticisms, and suggestions.

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The project would never have happened without the strong support and encouragement from CRC Press, including Ruijun He and his colleagues. We also thank them for their help in getting the book into its final form.



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# Comments and Errata

It seems that errors or typos in this book are inevitable. Any errors that remain are solely the author's responsibility. I am always more than happy to receive feedback on this book, especially constructive comments on how to improve this book. You can email your comments and suggestions to `zfcdo@cs.sjtu.edu.cn`.

*Shanghai, May, 2012.*

*Zhenfu Cao*

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# Chapter 1

## Introduction

In this chapter, we briefly review some security problems arisen in the network environments and present the main idea about how to use modern cryptographic techniques to solve these problems. We want to emphasize that cryptography is the building block of most solutions to security problems.

### 1.1 Trust Problem

#### 1.1.1 Trusted Domains Transfer Problem

We first discuss what the trusted domains transfer problem is. Suppose Alice and Bob belong to two trusted domains  $CA_1$  and  $CA_2$ , respectively, and they want to build trust relationship. In the public key infrastructure (PKI),  $CA_1$  and  $CA_2$  are two certificate authorities. Every user in a trusted domain has a public key certificate, which is a signature signed by the certificate authority to bind a public key with an identity. A certificate can be used to verify that a public key belongs to an individual.

However, Alice can only verify and trust the certificates from the certificate authority  $CA_1$ , and so can Bob from the certificate authority  $CA_2$ . In the scenario as shown in

Figure 1.1, how to build trust relationship between Alice in the trusted domain  $CA_1$  and Bob in  $CA_2$  is a practical problem.

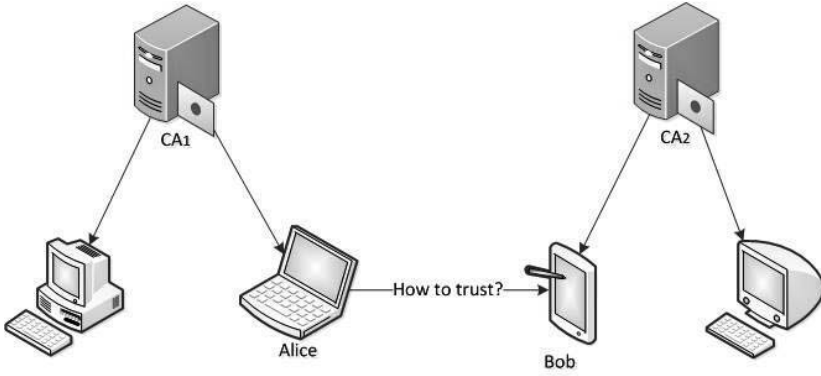


Figure 1.1: Trusted Domains Transfer Problem

**Solution:** The main idea to solve this problem is to set a transfer server, called proxy that is allowed to transform certificates from  $CA_1$  to  $CA_2$ , as shown in Figure 1.2. However, the proxy cannot generate new certificates in  $CA_1$  or  $CA_2$  by itself. We may require extra abilities of the proxy in some concrete applications.

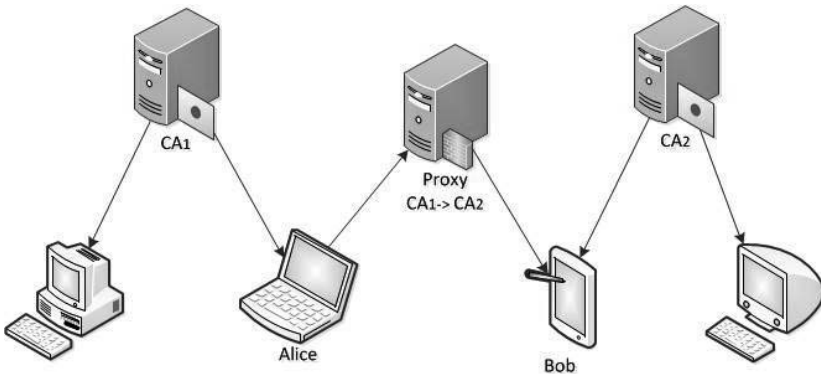


Figure 1.2: A Solution to Trusted Domains Transfer Problem

Sometimes it is desired that certificates in  $CA_1$  can be transformed into ones in  $CA_2$ , while certificates in  $CA_2$  cannot be transformed into ones in  $CA_1$ , as shown in the

Figure 1.3. This requires that the proxy can only do the *unidirectional* transformation. This new ability of proxy required in this case is authorized only by  $CA_2$ . If the ability of proxy is authorized by both  $CA_1$  and  $CA_2$ , then the proxy could do the *bidirectional* transformation.

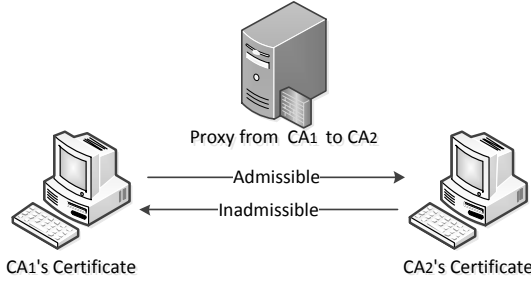


Figure 1.3: Unidirectional Transfer

On the other hand, the requirements in the trusted domains transfer problem could be further extended. As shown in Figure 1.4, certificates in  $CA_1$  can be transformed into the ones in  $CA_2$  via  $Proxy_1$ , and further transformed into the ones in  $CA_3$  via  $Proxy_2$ . If the process can continue multiple times, the method of trusted domains' transfer is *multiuse*. Otherwise, it is *single-use*.

Fortunately, we have a new cryptographic primitive called proxy re-signature [29] to solve the above problem. We will give a comprehensive introduction of proxy re-signature in Chapter 2 and discuss more applications of it in Chapter 6.

### 1.1.2 Trusted Server Problem

Cloud computing is drawing more and more attention from the information and communication technology community, since it can significantly reduce the costs of hardware and software resources in computing infrastructure.

In cloud computing, the cloud storage server is responsible for storing users' data

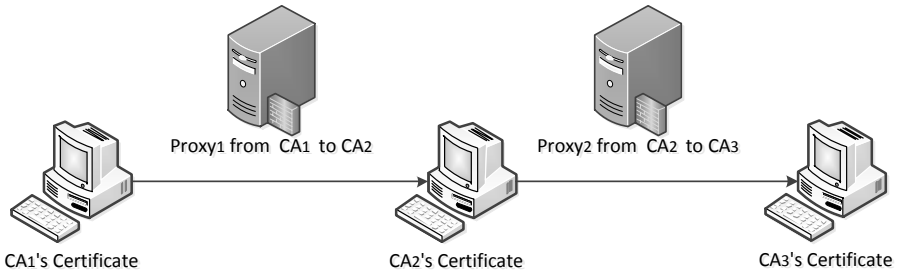


Figure 1.4: Multiuser Transfer.

including the sensitive data, and the cloud access control server is responsible for exerting control over who can access the data stored in the cloud storage server. It is usually required that the cloud access control server is fully trusted.

However, this requirement cannot be met in practice for two reasons. One is that the provider(s) of cloud access control service cannot be assumed to be fully trusted, because that he/she could become corrupted in some situations. The other is that intruders could break the cloud access control server even if the provider(s) is absolutely trusted. Hence, trusted server problem in cloud computing should be solved to put forward the development of cloud computing.

A possible solution is to store the encrypted plaintexts (i.e., ciphertexts) in the cloud storage server. If the ciphertexts are merely used by the encryptor himself/herself, the trusted server problem is easily solved. When the ciphertexts need to be shared by others and the access control server has no right to perform decryption, it is a challenging problem. Under this situation, we can conceive a following solution: let the encryptor authorize the access control server the right to transform the ciphertexts so that the delegated users can decrypt the resulting ciphertexts, but the access control server cannot decrypt the ciphertexts. This paradigm is shown in Figure 1.5.

The above conceived solution could be implemented if the access control server un-

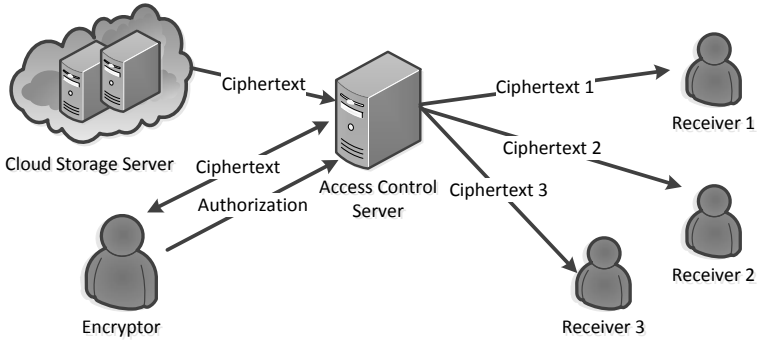


Figure 1.5: A Solution to Trusted Server Problem

der the authorization of the encryptor can transform the ciphertexts stored in the cloud storage server into a new form with the same plaintexts that can only be decrypted by the designated receivers. If we regard the access control server, encryptor, designated receivers, and authorization messages as the proxy, delegator, delegates, and re-encryption keys, respectively, the above solution becomes a particular case of proxy re-encryption.

Proxy re-encryption was proposed by Blaze et al. [29] and has many applications. According to the concrete applications, proxy re-encryption should satisfy other properties. We will discuss them in Chapters 2 and 6.

## 1.2 Ciphertext Access Control Problem

Assume that data owner intends to store a private message that is accessed by a specific set of users in a storage server. The current solution is that the data owner stores the data in plaintext form in the storage server, and the user's access rights are specified by access control lists that are created by the data owner and performed by the access control server as shown in Figure 1.6. The users specified by the access control



lists and verified by the access control server, can access the message. However, as mentioned in the above section, in practice, trust and security issues of the servers are always serious.

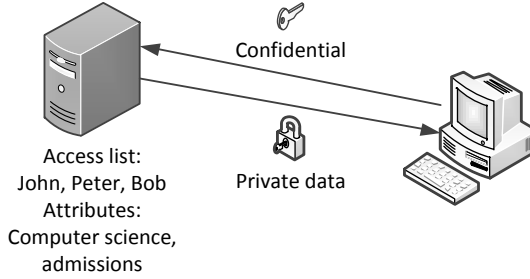


Figure 1.6: Access Control Model

A trivial method would be to store the data in ciphertext form in the servers. However current encryption systems cannot allow the ciphertext to be efficiently shared among a group of users. It becomes urgent to develop an efficient and flexible method to share data directly based on ciphertexts, which includes the access control policy. Fortunately, Bethencourt et al. [27] proposed such a cryptographic primitive called ciphertext-policy attribute-based encryption (CP-ABE), which initiates a new direction in solving the ciphertext access control problem.

The concept of attribute-based encryption (ABE) was first proposed by Goyal et al. [128] in 2005. They also constructed a concrete scheme called key-policy attribute-based encryption (KP-ABE). In KP-ABE as shown in Figure 1.7, each ciphertext is labeled by the encryptor with a set of descriptive attributes, and each private key is associated with an access structure that specifies which type of ciphertexts the key can decrypt. While in CP-ABE as shown in Figure 1.8, each private key is associated with a set of attributes and each ciphertext is associated with an access structure.

For example in Figure 1.8, the access structure of ciphertext File 1 is “{(Computer

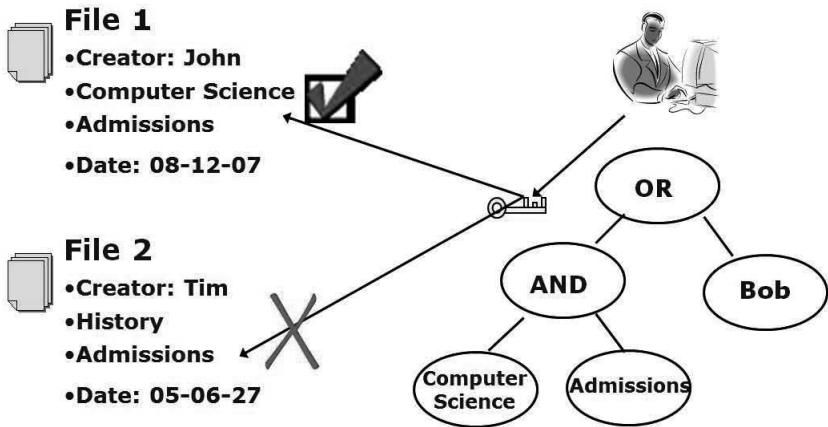


Figure 1.7: Key-Policy ABE

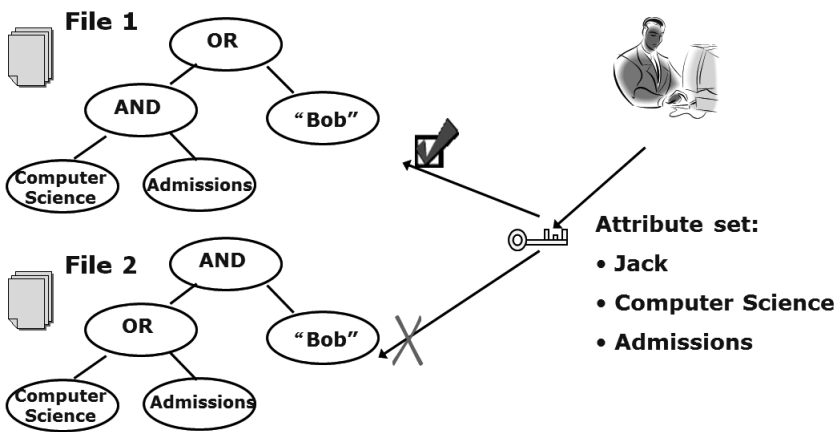


Figure 1.8: Ciphertext-Policy ABE

science) AND (Admissions)} OR (Bob)” representing that the qualified user should be an enrolled student in department of computer science or Bob. As Jack’s attribute set {Jack, Computer science, Admissions} satisfies the ciphertext access structure, he can decrypt File 1 successfully by using his private key. Moreover, he can not decrypt ciphertext File 2 since his attribute set does not satisfy the access structure “{(Computer science) OR (Admissions)} AND (Bob)” of File 2.

Recall the scenario discussed in Section 1.1.2, where the proxy can merely transform a ciphertext into another one that can only be decrypted by *a designated user*. But in practice, it is more usually required that the proxy is able to transform a ciphertext into another one that can be decrypted by *a group of designated users*. We propose a new concept of attribute-based proxy re-encryption [191], called ABPRE, which is a more powerful cryptographic primitive that can settle this problem efficiently.

To better understand the concept of ABPRE, we demonstrate an application scenario of personal information system in a university. In this system, there are some confidential records of grades of every student. These records are encrypted into a ciphertext under the access structure “((AGE > 40) AND (Tenure)).” Professors who are older than 40 and have a tenure position are qualified to retrieve the confidential records by using their own *different private keys*. Nevertheless, when these professors are on vacation, it is necessary to find some trustworthy delegates who are able to decrypt the ciphertext in time. Therefore, ABPRE allows a qualified professor to authorize a proxy (administrator) who can transform a ciphertext into another ciphertext encrypted with a different access structure so that the corresponding delegates can retrieve the records. For example, the delegated access structure can be defined by “(Secretary) AND (EXP ≥ 2)” that represents secretaries with at least 2 years working experience. Therefore, even if no qualified professor is available, some experienced secretaries can open the confidential records with the help of the authorized administrator.

A more general relationship between users and ciphertexts is shown in Figure 1.9.

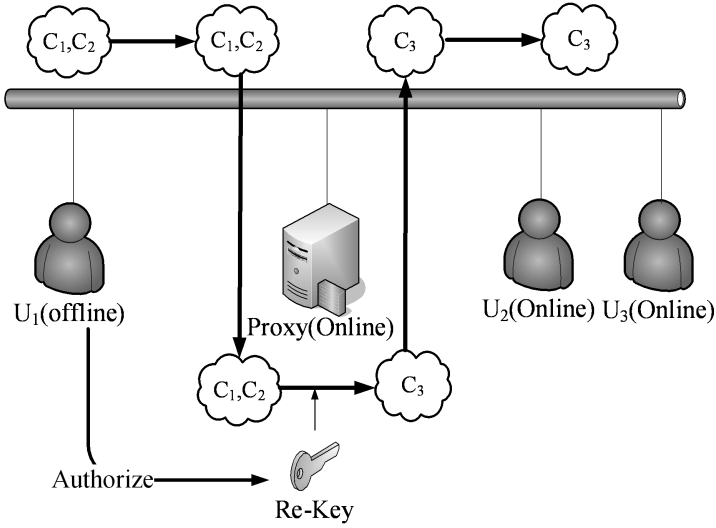


Figure 1.9: Attribute Based Proxy Re-encryption with Delegating Capabilities

Suppose there are three users and three ciphertext sets in this system, where user  $U_1$  is able to decrypt any ciphertext in sets  $C_1$  and  $C_2$  encrypted under access structures  $AS_1$  and  $AS_2$ , while users  $U_2$  and  $U_3$  are able to decrypt ciphertexts in  $C_3$  corresponding with  $AS_3$ . Then,  $U_1$  authorizes a proxy with a re-key that can be used to transform the ciphertext of  $C_1$  and  $C_2$  into that of  $C_3$ . In this way, even if  $U_1$  is offline,  $U_2$  and  $U_3$  could still retrieve the information encrypted in  $C_1$  and  $C_2$  with the help of  $U_1$ 's proxy.

### 1.3 Efficiency Problems in Multi-Message Cryptology

With the development of networks, there are many scenarios related to multiple requests processing as shown in Figure 1.10. The server receives multiple encrypted requests from different clients using the server's public key. In this scenario, is it necessary for the server to decrypt all the requests one by one? Generally, it is. But it is very

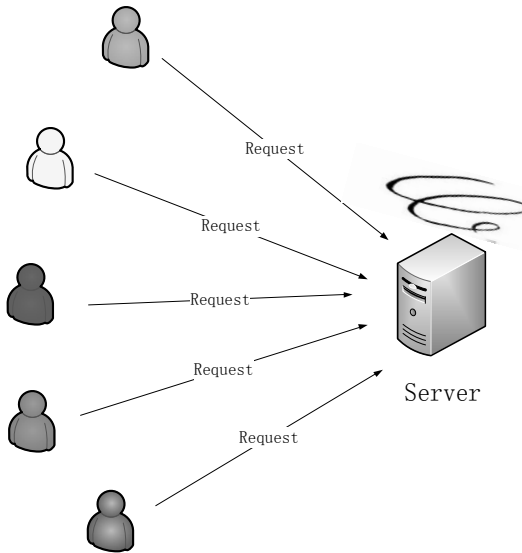


Figure 1.10: Multiple Requests Processing

inefficient when the number of requests becomes large. A similar scenario appears in the key exchange among the many clients and the central server.

Another scenario is that a famous star signs his books or discs for selling as shown in Figure 1.11. He has to meet a lot of his fans and sign his name many times, which leaves him exhausted. Digital signature continuously faces similar scenarios in the e-government or e-business. In these scenarios, is it necessary for them to sign one by one? Generally, they can sign only one by one, which results in low efficiency when the number of files to be signed is large.

A similar scenario exists in the current broadcast/multicast authentication, where the signatures need to be verified one by one. In this case, is it necessary to verify in such an inefficient way? Generally, it is. However, verifiers become exhausted due to excessive signature verification. The situation goes even worse in the case of the energy-constrained networks such as wireless sensor networks and mobile ad hoc networks.

Taking into account the above scenarios, it has become critical to introduce a new

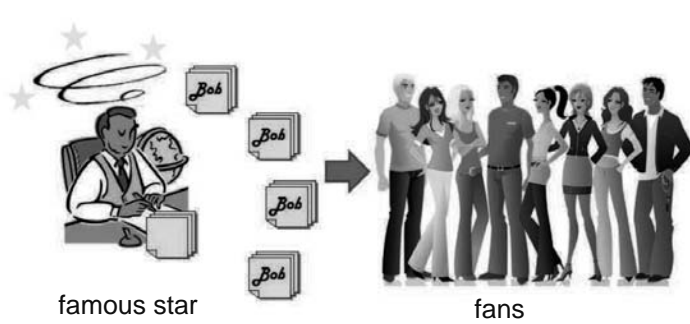


Figure 1.11: Signing Multiple Files

direction of modern cryptography—batch cryptography, which includes batch decryption, batch key agreement, and batch verification to solve the multi-message processing and their security and efficiency problems in real world. Readers can refer to Chapters 4 and 6 for more details.

# 1.4 The Challenges from Quantum and Biological Computing

In 1994, Shor [257] proposed a polynomial time algorithm for solving integer factorization problem <sup>1</sup> and the discrete logarithm problem using quantum computers. Meanwhile, most of modern cryptographic schemes are based on the hardness of integer factorization problem and discrete logarithm problem. If quantum computers become available, the modern cryptography based on the above two problems should be collapsed.

Biological computing, which refers to DNA computing [2] at present, is considered with super computing capability possibly more than quantum computing. If a biological computer appears, we believe that it will bring the same challenges to modern cryptog-

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<sup>1</sup>For the detail remarks of Shor’s algorithm, please refer the footnote on page 234 of this book.

raphy.

Therefore, what is the future cryptography? To answer this question, we suggest that one gets an insight into the history of cryptography.

Historically, stream cipher, block cipher, and related work use the same key for both encryption and decryption, which form the symmetric cryptography. The milestones in symmetric cryptography are communication theory of secrecy systems proposed by Shannon [243, 244] in 1949 and data encryption standard (DES) [222] in 1975.

In 1976, Diffie and Hellman [99] proposed the new direction of cryptography — asymmetric encryption in the network environment. Two different keys are required in their work: one is public for encryption and the other is private for decryption.

Although they did not propose any concrete encryption schemes, it paved the way for asymmetric encryption. After that, many researchers have proposed some concrete schemes such as RSA [233] based on integer factorization problem and ElGamal [107] based on discrete logarithm problem. Following this method, researchers have obtained many fruitful results, which form the asymmetric cryptography.

Both the symmetric cryptography and asymmetric cryptography are based on commutative algebraic structures. We call these kinds of cryptography commutative cryptography. As we have seen the transition of cryptography from “symmetric” to “asymmetric,” the next transition of cryptography might be from “commutative” to “noncommutative.” We refer to the cryptography based on the noncommutative algebraic structures as noncommutative cryptography.

Some computing hard problems based on noncommutative algebraic structures might be essential to resist quantum computing or biological computing. Thus, one possible direction of modern cryptography is to securely bridge commutative cryptography to noncommutative cryptography. Another possible direction is to explore noncommutative algebraic structures from the quantum theory and biological technology. We will discuss this detailedly in Chapters 5 and 6.

## 1.5 Organization

The theme of this book is to introduce some new directions of modern cryptography. The main idea we want to express is: cryptography is brought into birth and driven by application requirements. We will cover four kinds of modern cryptography: proxy re-cryptography, attribute-based cryptography, batch cryptography, and noncommutative cryptography. This book consists of the following Chapters. Chapter 1 introduces the background and motivation to write this book. Chapters 2, 3, 4, and 5, respectively, introduce proxy re-cryptography, attribute-based cryptography, batch cryptography, and noncommutative cryptography, including the fundamental definitions, security models, concrete schemes and security proof. In Chapter 6, we present some applications and challenging problems related to the above cryptography.



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# Chapter 2

## Proxy Re-Cryptography

### 2.1 Introduction

Proxy re-cryptography can be used to solve the trust problem in the scenarios mentioned in Chapter 1. Besides, there are many promising applications.

Many companies have developed digital rights management (DRM) technologies, which can prevent illegal redistribution of digital content. With DRM systems, the digital content can only be played in a specified device (regime). For example, a song playable in device (regime) *A* cannot be played in device (regime) *B*. However, it is reported that 86% of the consumers prefer to pay twice the price for a song that can run on any device than that with one single device [155]. Most of current interoperability architectures require to change the existing DRM systems significantly [176]; this modification cannot be adopted due to business reasons. At ACM DRM 2006, Taban *et al.* [271] proposed a new interoperability architecture, which does not change the existing DRM systems too much but maintains the DRM systems' security. In their architecture, only a new module called domain interoperability manager (DIM) is introduced. DIM applies a single signature scheme and a single public key encryption

scheme that can transform licences and content in regime  $A$  into ones in regime  $B$ , but it cannot generate valid licenses or content either in regime  $A$  or in regime  $B$ . It is easy to see that the traditional signature and public key encryption cannot support transformation, which, however, can be easily implemented by proxy re-cryptography.

In this chapter, we will give a comprehensive introduction of proxy re-cryptography, especially the security models of proxy re-cryptography. Furthermore, we introduce one proxy re-signature scheme and one proxy re-encryption scheme for illustration.

## 2.2 Proxy Re-Signature

Proxy re-signature (PRS), introduced by Blaze et al. [29] at Eurocrypt 1998, and formalized by Ateniese and Hohenberger [11] at ACM CCS 2005, allows a semi-trusted proxy to transform a delegatee's (Alice) signature into a delegator's (Bob) signature on the same message by using some additional information (a.k.a. re-signature key). The proxy, however, cannot generate arbitrary signatures on behalf of either the delegatee or the delegator.

### 2.2.1 Properties and Definition

Before giving the formal definition of PRS, we would like to give the desired properties of PRS [11].

- *Unidirectional*: In this scheme, a re-signature key allows the proxy to transform Alice's signature to Bob's but not vice versa. On the other hand, in a bidirectional scheme, the re-signature key allows the proxy to transform Alice's signature to Bob's as well as Bob's signature to Alice's.
- *Multiuse*: In a multiuse scheme, a transformed signature can be re-transformed again by a proxy. While in a single-use scheme, a proxy can transform only the

signatures not yet transformed.

- *Private proxy*: A proxy can keep the re-signature key as a secret in a private proxy scheme, but *anyone* can recompute the re-signature key by observing the re-signature process passively in a public proxy scheme.
- *Transparent*: In a transparent scheme, users may not even know the existence of a proxy.
- *Key-optimal*: In a key-optimal scheme, a user is required to protect and store only a small constant amount of secrets no matter how many signature delegations the user gives or accepts.
- *Noninteractive*: The delegatee is not required to participate in the delegation process. Bidirectional PRS cannot be noninteractive, since a delegator is also a delegatee.
- *Non-transitive*: A re-signature key cannot be generated from two other re-signature keys. For example, the re-signature key from Alice to Bob cannot be generated from the re-signature keys from Alice to Tina and Tina to Bob.
- *Temporary*: A re-signing right is temporary. This can be done by either revoking the right [11] or expiring the right.
- *Collusion-resistant*: The delegator can delegate the signing rights to the delegatee via the proxy, while keeping decryption rights for the same public key.

Now we give the formal definition of PRS.

**Definition 2.2.1** (Proxy Re-Signature). A PRS scheme consists of the following five probabilistic polynomial time (p.p.t.) algorithms:

- **KeyGen**: It takes as input the security parameter  $\lambda$ , and returns a verification key  $pk$  and a signing key  $sk$ .

- **ReKeyGen**: It takes as input delegatee Alice's key pair  $(pk_A, sk_A)$  and delegator Bob's key pair  $(pk_B, sk_B)$ , and returns a re-signature key  $rk_{A \rightarrow B}$  for the proxy. If the PRS scheme is unidirectional, the delegatee's signing key is not included in the input. If the PRS is bidirectional, the proxy can easily obtain  $rk_{B \rightarrow A}$  from  $rk_{A \rightarrow B}$ . In many bidirectional schemes [11, 29, 250],  $rk_{A \rightarrow B} = 1/rk_{B \rightarrow A}$ .
- **Sign**: It takes as input a signing key  $sk$ , a positive integer  $\ell$ , and a message  $m$  from the message space, and returns a signature  $\sigma$  at level  $\ell$ . If the PRS scheme is single-use, then  $\ell \in \{1, 2\}$ .
- **ReSign**: It takes as input a re-signature key  $rk_{A \rightarrow B}$  and a signature  $\sigma_A$  on a message  $m$  under  $pk_A$  at level  $\ell$ , and returns the signature  $\sigma_B$  on the same message  $m$  under  $pk_B$  at level  $\ell + 1$  if  $\text{Verify}(pk_A, m, \sigma_A, \ell) = 1$ , or reject otherwise. If the PRS scheme is single-use, then  $\ell = 1$ .
- **Verify**: It takes as input a verification key  $pk$ , a message  $m$  from the message space, a signature  $\sigma$ , and a positive integer  $\ell$ , and returns 1 if  $\sigma$  is a valid signature under  $pk$  at level  $\ell$  or 0 otherwise.

**Correctness.** The following property must be satisfied for the correctness of a PRS scheme: For any message  $m$  in the message space and any two key pairs  $(pk_A, sk_A)$  and  $(pk_B, sk_B)$ , let  $rk_{A \rightarrow B} \leftarrow \text{ReKeyGen}(pk_A, sk_A, pk_B, sk_B)$ , the following two equalities must hold:

$$\text{Verify}(pk_A, m, \sigma_A, \ell) = 1,$$

where  $\sigma_A$  is a signature on message  $m$  under  $pk_A$  at level  $\ell$  from **Sign**. If the PRS scheme is single-use, then  $\ell \in \{1, 2\}$ , or  $\ell \geq 1$  otherwise.

$$\text{Verify}(pk_B, m, \text{ReSign}(rk_{A \rightarrow B}, pk_A, m, \sigma'_A, \ell - 1), \ell) = 1.$$

If the PRS scheme is single-use,  $\sigma'_A$  is a signature on message  $m$  under  $pk_A$  from  $\text{Sign}$  with  $\ell = 2$ ; if it is multiuse,  $\sigma'_A$  could be a signature on message  $m$  under  $pk_A$  from  $\text{Sign}$  with  $\ell = 2$  or from  $\text{ReSign}$  with  $\ell > 2$ .

*Remark 2.2.2 (Two Types of Signatures).* In all existing unidirectional PRS schemes, a signature manifests in two types: the *owner-type* (i.e., the first-level defined in [11], and  $\ell = 1$  in this chapter) and the *nonowner-type* (i.e., the second-level signatures [11], and  $\ell > 1$  in this chapter). An *owner-type* signature can be computed only by the owner of the signing key via  $\text{Sign}$ , while a *nonowner-type* signature can be computed not only by the owner of the signing key via  $\text{Sign}$ , but also by collaboration between his/her proxy and delegatee via  $\text{ReSign}$ .

If there is only one signature type in a PRS scheme, the parameter  $\ell$  in all algorithms can be omitted.

### 2.2.2 Related Work

Though PRS has many applications as we mentioned in Chapter 1, it has a rather simple history. The first PRS scheme, which is multiuse, public proxy and bidirectional, was proposed by Blaze et al. [29] at Eurocrypt 1998. However, there was no follow-up until the work published by Ateniese and Hohenberger [11] at ACM CCS 2005. One of the reasons is that the definition of proxy re-signatures [29] was informal and could be easily confused with other signature variations. Ateniese and Hohenberger [11] first formalized the definition of security for PRS, referred as the AH model in this chapter, and then proposed three PRS schemes with security proofs. The first one is multiuse, private proxy and bidirectional; the second one is single-use, public proxy and unidirectional; and the third one is single-use, private proxy and unidirectional. Later, by using Waters' identity-based signature [289], we [250] proposed a multiuse, private proxy and bidirectional PRS scheme, and Chow and Phan [85] proposed a single-use, private proxy and unidirectional PRS scheme. Libert and Vergnaud [193] proposed a multiuse,

private proxy and unidirectional scheme. All the PRS schemes [11, 85, 193, 250] are proven secure in the AH model.

Recently, we [251] found that the AH model is suitable for almost all proxy re-signatures except the private proxy and unidirectional PRS. To deal with this problem, we [251] proposed an improvement of the AH model ( $AH^+$  model for short). Fortunately, the previous schemes proven secure in the AH model can still be proven secure in the  $AH^+$  model. Following this, we [252] further extended the  $AH^+$  model to the ID-based setting, and proposed a unidirectional, single-use, private proxy, and ID-based PRS.

Some PRS schemes in terms of the satisfied properties are summarized in Table 2.1.

Table 2.1: Properties of some PRS schemes

Property <sup>a</sup>	P1	P2	P3	P4	P5	P6	P7	P8	P9
BBS [29]	✗	✓	✗	✓	✓	✗	✗	✓	✗
AH05a [11]	✗	✓	✓	✓	✓	✗	✗	✓	✗
AH05b [11]	✓	✗	✗	✓	✓	✓	✓	✓	✓
AH05c [11]	✓	✗	✓	✓	✓	✓	✓	✓	✓
SCWL07a [250]	✗	✓	✓	✓	✓	✗	✗	✓	✗
SCWL07b [250]	✗	✓	✓	✓	✓	✗	✗	✓	✗
CP08 [85]	✓	✗	✓	✓	✓	✓	✓	✓	✓
LV08 [193]	✓	✓	✓	✓	✓	✓	✓	✓	✓
SWLX11 [252]	✓	✓	✓	✓	✓	✓	✓	✓	✓

<sup>a</sup> P1, ..., P9 denote unidirectional, multiuse, private proxy, transparent, key-optimal, non-transitive, temporary, and collusion-resistant, respectively.

In the rest of this section, we first introduce the security model of PRS—the AH model. Then we introduce the multiuse, private proxy and bidirectional PRS scheme [250] proposed at Indocrypt 2007. In what follows, we show the incompleteness of the AH model, and give an improvement,  $AH^+$  model [251].

### 2.2.3 Security Model: The AH Model

The AH model mainly deals with the unforgeability of signatures; it contains external security and internal security. These two parts contain four security games between an

adversary  $\mathcal{A}$  and a challenger  $\mathcal{C}$ . Before the games start, the adversary should decide which users (delegatee/delegator) are to be corrupted. Furthermore, all the verification keys in the games are generated by the challenger. If the adversary obtains the signing key of a user via  $\mathcal{O}_{sk}$ , then the user is declared corrupted. We assume that the adversary never invokes a query twice. If so, the challenger simply returns the previous value.

The following oracles could be queried in the four games:

- *Verification key query  $\mathcal{O}_{pk}$* : On receiving an index  $i$  from the adversary  $\mathcal{A}$ , the challenger  $\mathcal{C}$  responds by running  $\text{KeyGen}(1^\lambda)$  to get a key pair  $(pk_i, sk_i)$ , and forwards the verification key  $pk_i$  to the adversary. Finally, the challenger records  $(pk_i, sk_i)$  in Table  $T_k$ , which is initialized as empty and used to record the key pairs of users. In the following oracles,  $sk_i$  denotes as the signing key corresponding to the verification key  $pk_i$ .
- *Signing key query  $\mathcal{O}_{sk}$* : On receiving a verification key  $pk_i$  from the adversary  $\mathcal{A}$ , the challenger  $\mathcal{C}$  responds with  $sk_i$  which is the associated value with  $pk_i$  in Table  $T_k$ , if  $pk_i$  is corrupted; otherwise the challenger  $\mathcal{C}$  responds with **reject**.
- *Re-signature key query  $\mathcal{O}_{rk}$* : On receiving two verification keys  $pk_i, pk_j$  ( $pk_i \neq pk_j$ ) from the adversary  $\mathcal{A}$ , the challenger  $\mathcal{C}$  responds with  $\text{ReKeyGen}(pk_i, sk_i, pk_j, sk_j)$ . Note that for bidirectional PRS, we consider queries with  $(pk_i, pk_j)$  and  $(pk_j, pk_i)$  are identical.
- *Signature query  $\mathcal{O}_s$* : On receiving a verification key  $pk_i$  and a message  $m_i$  from the adversary  $\mathcal{A}$ , the challenger  $\mathcal{C}$  responds with  $\text{Sign}(sk_i, m, 1)$ .
- *Re-signature query  $\mathcal{O}_{rs}$* : On receiving two verification keys  $pk_i, pk_j$  ( $pk_i \neq pk_j$ ), a message  $m_i$ , and a signature  $\sigma_i$  at level  $\ell$  from the adversary  $\mathcal{A}$ , the challenger  $\mathcal{C}$  responds with  $\text{ReSign}(\text{ReKeyGen}(pk_i, sk_i, pk_j, sk_j), pk_i, m_i, \sigma_i, \ell)$ .

Now, we introduce the four security games as follows:



**External Security:** This security protects a user from outside adversaries other than the proxy and any delegation parties. The game goes as follows:

- *Queries:* The adversary  $\mathcal{A}$  can make queries to oracles  $\mathcal{O}_{pk}$ ,  $\mathcal{O}_{rs}$ , and  $\mathcal{O}_s$  adaptively.
- *Forgery:* The adversary  $\mathcal{A}$  outputs a signature  $\sigma^*$  on message  $m^*$  on behalf of  $pk^*$  at level  $\ell^*$ . The adversary is declared the winner of the game, if all of the following requirements are satisfied:
  - $\text{Verify}(pk^*, m^*, \sigma^*, \ell^*) = 1$ ;
  - The adversary  $\mathcal{A}$  never obtains a signature on a message  $m^*$  under the verification key  $pk^*$  by querying  $\mathcal{O}_s$  with  $(pk^*, m^*)$ , or by querying  $\mathcal{O}_{rs}$  with  $(\cdot, pk^*, m^*, \cdot, \cdot)$ .

A PRS scheme has external security if and only if for security parameter  $\lambda$  and all p.p.t. algorithms  $\mathcal{A}$ ,  $\Pr[\mathcal{A} \text{ wins}]$  is *negligible*.

**Internal Security:** This security protects a user from inside adversaries who can be any parties, i.e., the proxy, the delegatee, or the delegator, in a PRS scheme. It can be classified into the following three types:

**Limited Proxy:** In this case, only the proxy is a potential adversary  $\mathcal{A}$ . We must be sure that the proxy cannot produce signatures on behalf of either the delegator or the delegatee except the signatures produced by the delegatee and delegated to the proxy to re-sign. The game goes as follows:

- *Queries:* The adversary  $\mathcal{A}$  can make queries to oracles  $\mathcal{O}_{pk}$ ,  $\mathcal{O}_{rk}$ ,  $\mathcal{O}_{rs}$ , and  $\mathcal{O}_s$  adaptively.
- *Forgery:* The adversary  $\mathcal{A}$  outputs a signature  $\sigma^*$  on message  $m^*$  on behalf of  $pk^*$  at level  $\ell^*$ . The adversary is declared the winner of the game, if all of the

following requirements are satisfied:

- $\text{Verify}(pk^*, m^*, \sigma^*, \ell^*) = 1$ ;
- The adversary  $\mathcal{A}$  never obtains a signature on message  $m^*$  under the verification key  $pk^*$  by querying  $\mathcal{O}_s$  with  $(pk^*, m^*)$ , or by querying  $\mathcal{O}_{rs}$  with  $(\cdot, pk^*, m^*, \cdot, \cdot)$ ;
- If the PRS scheme is single-use, and the adversary  $\mathcal{A}$  obtains a re-signature key from  $pk_i$  to  $pk^*$  directly from  $\mathcal{O}_{rk}$ , then the adversary  $\mathcal{A}$  never obtains a signature on message  $m^*$  under the verification key  $pk_i$  by querying  $\mathcal{O}_s$  with  $(pk_i, m^*)$ ;
- If the PRS scheme is multiuse, and the adversary  $\mathcal{A}$  obtains re-signature keys corresponding to  $(pk_{i_1}, pk_{i_2})$ ,  $(pk_{i_2}, pk_{i_3})$ ,  $\dots$ ,  $(pk_{i_t}, pk^*)$ , then the adversary  $\mathcal{A}$  never obtains a signature on message  $m^*$  under the verification key  $pk_i$  ( $i \in \{i_1, \dots, i_t\}$ ) by querying  $\mathcal{O}_s$  with  $(pk_i, m^*)$ , or by querying  $\mathcal{O}_{rs}$  with  $(\cdot, pk_i, m^*, \cdot, \cdot)$ .

A PRS scheme has limited proxy security if and only if for security parameter  $\lambda$  and all p.p.t. algorithms  $\mathcal{A}$ ,  $\Pr[\mathcal{A} \text{ wins}]$  is *negligible*.

**Delegatee Security:** In this case, the proxy and delegator may collude with each other. This security guarantees that their collusion cannot produce any signatures on behalf of the delegatee.

The game goes as follows:

- *Queries:* The adversary  $\mathcal{A}$  can make queries to oracles  $\mathcal{O}_{pk}$ ,  $\mathcal{O}_{sk}$ ,  $\mathcal{O}_{rk}$ ,  $\mathcal{O}_{rs}$ , and  $\mathcal{O}_s$  adaptively.
- *Forgery:* The adversary  $\mathcal{A}$  outputs a signature  $\sigma^*$  on message  $m^*$  on behalf of  $pk^*$  at level  $\ell^*$ . The adversary is declared the winner of the game, if all of the following requirements are satisfied:

- $pk^*$  is uncorrupted;
- $pk^*$  has not been a delegator in this game;
- $\text{Verify}(pk^*, m^*, \sigma^*, \ell^*) = 1$ ;
- The adversary  $\mathcal{A}$  never obtains a signature on message  $m^*$  under the verification key  $pk^*$  by querying  $\mathcal{O}_s$  with  $(pk^*, m^*)$ .

A PRS scheme has delegatee security if and only if for security parameter  $\lambda$  and all p.p.t. algorithms  $\mathcal{A}$ ,  $\Pr[\mathcal{A} \text{ wins}]$  is *negligible*.

**Delegator Security:** In this case, the proxy and delegatee may collude with each other. This security guarantees that their collusion cannot produce any owner-type signatures on behalf of the delegator.

The game goes as follows:

- *Queries:* The adversary  $\mathcal{A}$  can make queries to oracles  $\mathcal{O}_{pk}$ ,  $\mathcal{O}_{sk}$ ,  $\mathcal{O}_{rk}$ ,  $\mathcal{O}_{rs}$ , and  $\mathcal{O}_s$  adaptively.
- *Forgery:* The adversary  $\mathcal{A}$  outputs a signature  $\sigma^*$  on message  $m^*$  on behalf of  $pk^*$  at the first level. The adversary is declared the winner of the game, if all of the following requirements are satisfied:
  - $pk^*$  is uncorrupted;
  - $pk^*$  has not been a delegatee in this game;
  - $\text{Verify}(pk^*, m^*, \sigma^*, 1) = 1$ .

A PRS scheme has delegator security if and only if for security parameter  $\lambda$  and all p.p.t. algorithms  $\mathcal{A}$ ,  $\Pr[\mathcal{A} \text{ wins}]$  is *negligible*.

*Remark 2.2.3.* Since a delegator is also a delegatee in bidirectional PRS, we do not consider the delegatee security and delegator security for bidirectional PRS with only one

signature type. While we have a weaker security for protecting the delegatee (delegator) in this kind of PRS, which is collusion resistance. It guarantees that the signing key of a delegatee (delegator) cannot be revealed by the collusion of the delegator (delegatee) and the proxy. On the other hand, it is easy to show that the delegatee security and delegator security imply the collusion resistance. In particular, if the adversary cannot generate a valid forgery, it definitely cannot obtain the corresponding signing key.

### 2.2.4 Multiuse, Private Proxy and Bidirectional Scheme

In this section, we [250] introduce the multiuse, private proxy and bidirectional PRS scheme (denoted as  $S_{mb}$ ) proposed at Indocrypt 2007. The security is proved in the standard model based on the computational Diffie–Hellman (CDH) assumption, that is, it is hard to compute  $g^{xy}$  given  $(g, g^x, g^y)$  where  $g$  is a random element in a cryptographic group  $\mathbb{G}$  with a prime order  $q$ , and  $x, y$  are random numbers in  $\mathbb{Z}_q^*$ .

The system parameter of scheme  $S_{mb}$  is  $(\mathbb{G}, \mathbb{G}_T, q, g, g_2, e, H_w)$ , where  $\mathbb{G}$  and  $\mathbb{G}_T$  are bilinear groups with prime order  $q$ ;  $g$  and  $g_2$  are two random elements of  $\mathbb{G}$ ;  $e$  is an admissible pairing,  $e : \mathbb{G} \times \mathbb{G} \rightarrow \mathbb{G}_T$ ; and  $H_w$  is a Waters' function [289],  $H_w(m) = u' \cdot \prod_{i \in \mathcal{U}} u_i$ , where  $m$  is an  $n_m$ -bit message,  $u', u_1, u_2, \dots, u_{n_m}$  are random elements in  $\mathbb{G}$ ,  $\mathcal{U} \subset \{1, \dots, n_m\}$  is the set of indices  $i$  such that  $m[i] = 1$ , and  $m[i]$  is the  $i$ -th bit of  $m$ .

- **KeyGen**: It outputs the key pair  $(pk, sk) = (g^a, a)$ , where  $a$  is a random number in  $\mathbb{Z}_q^*$ .
- **ReKeyGen**: On input two signing keys  $sk_A = a$  and  $sk_B = b$ , output the re-signature key  $rk_{A \rightarrow B} = b/a \bmod q$ . It is easy to see that  $rk_{B \rightarrow A} = 1/rk_{A \rightarrow B} = a/b \bmod q$ .

(Note that the re-signature key can be obtained by the following method [11]: (1) the proxy sends a random number  $r \in \mathbb{Z}_q^*$  to the delegatee Alice, (2) then Alice

sends  $r/a$  to the delegator Bob, (3) Bob sends  $rb/a$  to the proxy, (4) finally, the proxy obtains  $b/a$ . The communications between any two entities are via private and authenticated channels.)

- **Sign:** On input a signing key  $sk = a$  and an  $n_m$ -bit message  $m$ , output the signature  $\sigma = (\sigma_1, \sigma_2) = (g_2^a \cdot H_w(m)^r, g^r)$ , where  $r$  is chosen randomly from  $\mathbb{Z}_q^*$ .
- **ReSign:** On input a re-signature key  $rk_{A \rightarrow B}$ , a verification key  $pk_A$ , a signature  $\sigma_A = (\sigma_{A,1}, \sigma_{A,2})$ , and an  $n_m$ -bit message  $m$ , check that  $\text{Verify}(pk_A, m, \sigma_A) = 1$ . If  $\sigma_A$  is invalid, output **reject**; otherwise, choose a random number  $v \in \mathbb{Z}_q^*$  and output  $\sigma_B = (\sigma_{A,1}^{rk_{A \rightarrow B}} \cdot H_w(m)^v, \sigma_{A,2}^{rk_{A \rightarrow B}} \cdot g^v)$ .

Note that we have:

$$\begin{aligned} (\sigma_{A,1}^{rk_{A \rightarrow B}} \cdot H_w(m)^v, \sigma_{A,2}^{rk_{A \rightarrow B}} \cdot g^v) &= (g_2^b \cdot H_w(m)^{rb/a} H_w(m)^v, g^{rb/a} g^v) \\ &= (g_2^b H_w(m)^{r'}, g^{r'}) \end{aligned}$$

where  $r' = rb/a + v \bmod q$ .

- **Verify:** On input a verification key  $pk$ , an  $n_m$ -bit message  $m$ , and a signature  $\sigma = (\sigma_1, \sigma_2)$ , output 1, if  $e(pk, g_2)e(\sigma_2, H_w(m)) = e(\sigma_1, g)$ , or 0 otherwise.

**Theorem 2.2.4.** *In the standard model, the bidirectional PRS scheme  $S_{mb}$  is correct and existentially unforgeable under the CDH assumption in  $\mathbb{G}$ .*

*Proof.* The correctness property is easily observable.

To prepare the simulation, simulator  $\mathcal{B}$  first sets  $\ell_m = 2(q_s + q_{rs})$ , and randomly chooses a number  $k_m$ , such that  $0 \leq k_m \leq n_m$ , and  $\ell_m(n_m + 1) < q$ .  $\mathcal{B}$  then chooses  $n_m + 1$  random numbers  $x', x_i (i = 1, \dots, n_m)$  from  $\mathbb{Z}_{\ell_m}$ . Lastly,  $\mathcal{B}$  chooses  $n_m + 1$  random numbers  $y', y_i (i = 1, \dots, n_m)$  from  $\mathbb{Z}_q^*$ .  $q_s$  and  $q_{rs}$  are the number of queries to  $\mathcal{O}_s$  and  $\mathcal{O}_{rs}$ , respectively.

To make expression simpler, we use the following notations:

$$F(m) = x' + \sum_{i \in \mathcal{U}} x_i - \ell_m k_m \text{ and } J(m) = y' + \sum_{i \in \mathcal{U}} y_i.$$

Now,  $\mathcal{B}$  sets the public parameters:

$$g_2 = g^b, \quad u' = g_2^{x' - \ell_m k_m} g^{y'}, \quad u_i = g_2^{x_i} g^{y_i} \quad (1 \leq i \leq n_m).$$

Note that for any message  $m$ , there exists the following equality:

$$H_w(m) = u' \prod_{i \in \mathcal{U}} u_i = g_2^{F(m)} g^{J(m)}.$$

In the following, we focus on external security and internal security (limited proxy) of scheme  $S_{mb}$ .

### External Security:

- *Queries:*  $\mathcal{B}$  builds the following oracles:

- $\mathcal{O}_{pk}$ : On input index  $i$ ,  $\mathcal{B}$  chooses a random  $x_i \in \mathbb{Z}_q^*$ , and guess whether  $pk_i = pk^*$ . If it is, it outputs  $pk_i = (g^a)^{x_i}$ ; otherwise, it outputs  $pk_i = g^{x_i}$ . Finally,  $\mathcal{B}$  records  $(pk_i, x_i)$  in Table  $T_k$ .

- $\mathcal{O}_s$ : On input  $(pk_i, m)$ ,  $\mathcal{B}$  obtains  $(pk_i, x_i)$  from Table  $T_k$ . If  $pk_i \neq pk^*$ , it performs **Sign** with  $x_i$ ; otherwise, it performs as follows.

- \* If  $F(m) \neq 0 \pmod q$ ,  $\mathcal{B}$  picks a random  $r \in \mathbb{Z}_q^*$  and computes the signature as,

$$\begin{aligned} \sigma &= (((g^a)^{-J(m)/F(m)} (g^{J(m)} g_2^{F(m)r})^{x_i}, ((g^a)^{-1/F(m)} g^r)^{x_i}) \\ &= (\sigma_1, \sigma_2) \end{aligned}$$

For  $\tilde{r} = (r - a/F(m)) \cdot x_i$ , we have that

$$\begin{aligned}
 & ((g^a)^{-J(m)/F(m)} (g^{J(m)} g_2^{F(m)})^r)^{x_i} \\
 = & (g_2^a (g_2^{F(m)} g^{J(m)})^{-a/F(m)} (g^{J(m)} g_2^{F(m)})^r)^{x_i} \\
 = & (g_2^a (g_2^{F(m)} g^{J(m)})^{r-a/F(m)})^{x_i} \\
 = & g_2^{ax_i} H_w(m)^{\tilde{r}},
 \end{aligned}$$

and

$$\begin{aligned}
 ((g^a)^{-1/F(m)} g^r)^{x_i} &= (g^{r-a/F(m)})^{x_i} \\
 &= g^{\tilde{r}},
 \end{aligned}$$

which shows that  $\sigma$  has the correct signature as in the actual scheme.

\* If  $F(m) = 0 \pmod q$ ,  $\mathcal{B}$  is unable to compute the signature  $\sigma$  and outputs **failure**.

–  $\mathcal{O}_{rs}$ : On input  $(pk_i, pk_j, m, \sigma)$ . If  $\text{Verify}(pk_i, m, \sigma) \neq 1$ ,  $\mathcal{B}$  outputs **reject**. Otherwise,  $\mathcal{B}$  queries  $\mathcal{O}_s$  with  $(pk_j, m)$  and returns the resulting value.

- *Forgery*: If  $\mathcal{B}$  does not output **failure** in any query above,  $\mathcal{A}$  will, with probability at least  $\varepsilon$ , return a message  $m^*$  and a valid forgery  $\sigma^* = (\sigma_1^*, \sigma_2^*)$  on behalf of  $pk^*$ . It is easy to see that  $\mathcal{B}$  guesses the right  $pk^*$  with  $1/q_{pk}$  at least, where  $q_{pk}$  is the number of queries to  $\mathcal{O}_{pk}$ . If  $F(m^*) \neq 0 \pmod q$ ,  $\mathcal{B}$  outputs **failure**. Otherwise, the forgery must be of the form, for some  $r^* \in \mathbb{Z}_q^*$ ,

$$\begin{aligned}
 \sigma^* &= (g^{abx^*} (g_2^{F(m^*)} g^{J(m^*)})^{r^*}, g^{r^*}) \\
 &= (g^{abx^* + J(m^*)r^*}, g^{r^*}) \\
 &= (\sigma_1^*, \sigma_2^*).
 \end{aligned}$$

To solve the CDH instance,  $\mathcal{B}$  outputs  $((\sigma_1^*) \cdot (\sigma_2^*)^{-J(m^*)})^{1/x^*} = g^{ab}$ , where  $x^*$  is the corresponding value in Table  $T_k$  with  $pk^*$ .

At last, we need to bound the probability that  $\mathcal{B}$  completes the simulation without outputting `failure`. We require that all signature and re-signature queries on a message  $m$  along with  $pk^*$  have  $F(m) \neq 0 \bmod q$ , and that  $F(m^*) = 0 \bmod q$ . Let  $m_1, \dots, m_{q_Q}$  be the messages appearing in signature or re-signature queries not involving the message  $m^*$ . Clearly,  $q_Q \leq q_s + q_{rs}$ . We define the events  $E_i$ ,  $E'_i$ , and  $E^*$  as:

$$E_i : F(m_i) \neq 0 \bmod q, \quad E'_i : F(m_i) \neq 0 \bmod \ell_m, \quad E^* : F(m^*) = 0 \bmod q.$$

From  $\ell_m(n_m + 1) < q$  and  $x', x_i (i = 1, \dots, n_m) \in \mathbb{Z}_{\ell_m}$ , we have  $0 \leq l_m k_m < q$  and  $0 \leq x' + \sum_{i \in \mathcal{U}} x_i < q$ . Then  $F(m) = 0 \bmod q$  implies  $F(m) = 0 \bmod \ell_m$ . Hence,  $F(m) \neq 0 \bmod \ell_m$  implies  $F(m) \neq 0 \bmod q$ . Since  $k_m, x', x_i (i = 1, \dots, n_m)$  are chosen randomly, we have

$$\begin{aligned} \Pr[E^*] &= \Pr[F(m^*) = 0 \bmod q \wedge F(m^*) = 0 \bmod \ell_m] \\ &= \Pr[F(m^*) = 0 \bmod \ell_m] \Pr[F(m^*) = 0 \bmod q | F(m^*) = 0 \bmod \ell_m] \\ &= \frac{1}{\ell_m} \cdot \frac{1}{n_m + 1} \end{aligned}$$

and

$$\begin{aligned} \Pr[\bigwedge_{i=1}^{q_Q} E'_i | E^*] &= 1 - \Pr[\bigvee_{i=1}^{q_Q} \neg E'_i | E^*] \\ &\geq 1 - \sum_{i=1}^{q_Q} \Pr[E'_i | E^*] \\ &= 1 - \frac{q_s + q_{rs}}{\ell_m} \\ &= 1/2 \end{aligned}$$



The probability of  $\mathcal{B}$  not outputting failure is

$$\begin{aligned}
 \Pr[\neg \text{failure}] &\geq \Pr\left[\bigwedge_{i=1}^{q_Q} E_i \wedge E^*\right] \\
 &\geq \Pr\left[\bigwedge_{i=1}^{q_Q} E'_i \wedge E^*\right] \\
 &= \Pr[E^*] \Pr\left[\bigvee_{i=1}^{q_Q} \neg E'_i \mid E^*\right] \\
 &\geq \frac{1}{2\ell_m(n_m + 1)}
 \end{aligned}$$

**Internal Security:** Since a delegator in scheme  $S_{mb}$  is also a delegatee and only one signature type exists in scheme  $S_{mb}$ , we only consider limited proxy security here.

• *Queries:*  $\mathcal{B}$  builds the following oracles:

- $\mathcal{O}_{pk}$ : On input index  $i$ ,  $\mathcal{B}$  chooses a random  $x_i \in \mathbb{Z}_q^*$ , and guesses whether  $\mathcal{A}$  will issue the re-signature key queries with  $(pk_i, pk_{i_1}), (pk_{i_1}, pk_{i_2}), \dots, (pk_{i_j}, pk^*)$  or  $pk_i = pk^*$ . If it is,  $\mathcal{B}$  sets  $\theta_i = 0$  and outputs  $pk_i = (g^a)^{x_i}$ ; otherwise, it sets  $\theta_i = 1$  and outputs  $pk_i = g^{x_i}$ . Finally,  $\mathcal{B}$  records  $(pk_i, x_i, \theta_i)$  in Table  $T_k$ .
- $\mathcal{O}_s$ : On input  $(pk_i, m)$ ,  $\mathcal{B}$  obtains  $(pk_i, x_i, \theta_i)$  from Table  $T_k$ . If  $\theta_i = 1$ ,  $\mathcal{B}$  returns  $\text{Sign}(x_i, m)$ ; otherwise,  $\mathcal{B}$  proceeds as follows:

\* If  $F(m) \not\equiv 0 \pmod q$ ,  $\mathcal{B}$  picks a random  $r \in \mathbb{Z}_q^*$  and computes the signature as,

$$\begin{aligned}
 \sigma &= (((g^a)^{-J(m)/F(m)} (g^{J(m)} g_2^{F(m)r})^{x_i}, ((g^a)^{-1/F(m)} g^r)^{x_i}) \\
 &= (\sigma_1, \sigma_2).
 \end{aligned}$$

\* If  $F(m) \equiv 0 \pmod q$ ,  $\mathcal{B}$  is unable to compute the signature  $\sigma$  and outputs failure.

- $\mathcal{O}_{rk}$ : On input  $(pk_i, pk_j)$ , if  $\theta_i = \theta_j$ ,  $\mathcal{B}$  returns  $rk_{i \rightarrow j} = (x_j/x_i) \bmod q$ ; else,  $\mathcal{B}$  outputs **failure**.
- $\mathcal{O}_{rs}$ : On input  $(pk_i, pk_j, m, \sigma)$ . If  $\text{Verify}(pk_i, m, \sigma) \neq 1$ ,  $\mathcal{B}$  outputs **reject**; otherwise,  $\mathcal{B}$  queries  $\mathcal{O}_s$  with  $(pk_j, m)$  and returns the resulting value.
- *Forgery*: If  $\mathcal{B}$  does not output **failure** in any query above,  $\mathcal{A}$  will, with probability at least  $\varepsilon$ , return a message  $m^*$  and a valid forgery  $\sigma^* = (\sigma_1^*, \sigma_2^*)$  on behalf of  $pk^*$ . If  $F(m^*) \neq 0 \bmod q$  or  $\theta^* = 1$ ,  $\mathcal{B}$  outputs **failure**. Otherwise, the forgery must be of the form, for some  $r^* \in \mathbb{Z}_q^*$ ,

$$\begin{aligned}
\sigma^* &= (g^{abx^*} (g_2^{F(m^*)} g^{J(m^*)})^{r^*}, g^{r^*}) \\
&= (g^{abx^* + J(m^*)r^*}, g^{r^*}) \\
&= (\sigma_1^*, \sigma_2^*).
\end{aligned}$$

To solve the CDH instance,  $\mathcal{B}$  outputs  $((\sigma_1^*) \cdot (\sigma_2^*)^{-J(m^*)})^{1/x^*} = g^{ab}$ , where  $x^*$  is the corresponding value in Table  $T_k$  with  $pk^*$ .

At last, we need to bound the probability that  $\mathcal{B}$  completes the simulation without outputting **failure**. We require the followings:

**R1**  $\theta_i = \theta_j$  in  $\mathcal{O}_{rk}$ .

**R2**  $\theta^* = 0$ .

**R3**  $F(m) \neq 0 \bmod q$  with the corresponding  $\theta = 0$  in  $\mathcal{O}_s$  and  $\mathcal{O}_{rs}$ .

**R4**  $F(m^*) = 0 \bmod q$ .

It is easy to see that requirements R1 and R2 will be satisfied with the probability  $1/2^{q_{rk}}$  and  $1/q_{pk}$  at least, respectively, where  $q_{rk}$  is the number of queries to  $\mathcal{O}_{rk}$ .

Furthermore, we can use the similar analysis for external security to conclude that requirements R3 and R4 will be satisfied simultaneously with the probability  $\frac{1}{2\ell_m(n_m+1)}$  at least.

This completes the proof.  $\square$

### 2.2.5 Incompleteness of the AH Model

The AH model is designed for all kinds of PRS, and almost all existing PRS schemes are proven secure in the AH model. However, as we will see later, the AH model is not complete. In particular, the AH model is not suitable for the unidirectional and private proxy PRS. In this section, we propose a private proxy and unidirectional scheme, named  $S_{ins}$ , which is proven secure in the AH model, but it cannot provide all the required security properties. This fact shows that the AH model is not complete.

The public parameter of scheme  $S_{ins}$  is  $(q, g, \mathbb{G}, \mathbb{G}_T, e, H)$ , where  $\mathbb{G}$  and  $\mathbb{G}_T$  are bilinear groups with prime order  $q$ ,  $g$  is a random element of  $\mathbb{G}$ ,  $e$  is an admissible pairing,  $e : \mathbb{G} \times \mathbb{G} \rightarrow \mathbb{G}_T$ , and  $H : \{0, 1\}^* \rightarrow \mathbb{G}$  is a cryptographic hash function.

- **KeyGen:** It selects a random number  $a \in \mathbb{Z}_q^*$ , and outputs the key pair  $(pk, sk) = (g^a, a)$ .
- **ReKeyGen:** On input the delegatee's verification key  $pk_A = g^a$  and the delegator's signing key  $sk_B = b$ , it outputs the re-signature key

$$rk_{A \rightarrow B} = (rk_{A \rightarrow B}^{(1)}, rk_{A \rightarrow B}^{(2)}, rk_{A \rightarrow B}^{(3)}) = (r', (pk_A)^{r'}, H(g^{a \cdot r'} || 2)^{1/b}),$$

where  $r'$  is a random number in  $\mathbb{Z}_q^*$  determined by the delegator.

- **Sign:** On input a signing key  $sk = a$ , a message  $m$  in message space and an integer  $\ell \in \{1, 2\}$ ,

- if  $\ell = 1$ , it outputs an *owner-type* signature

$$\sigma = (A, B, C) = (H(m||0)^r, g^r, H(g^r||1)^a),$$

- if  $\ell = 2$ , it outputs a *nonowner-type* signature

$$\sigma = (A, B, C, D, E) = (H(m||0)^{r_1}, g^{r_1}, H(g^{r_1}||1)^{r_2}, g^{r_2}, H(g^{r_2}||2)^{1/a}).$$

- **ReSign:** Given an *owner-type* signature  $\sigma$  at level 1, a re-signature key  $rk_{A \rightarrow B} = (rk_{A \rightarrow B}^{(1)}, rk_{A \rightarrow B}^{(2)}, rk_{A \rightarrow B}^{(3)})$ , a verification key  $pk_A$ , and a message  $m$ , this algorithm first checks  $\text{Verify}(pk_A, m, \sigma, 1) = 1$ . If it does not hold, outputs **reject**; otherwise, outputs

$$\begin{aligned} \sigma' &= (A', & B', & C', & D', & E') \\ &= (A, & B, & C^{rk_{A \rightarrow B}^{(1)}}, & rk_{A \rightarrow B}^{(2)}, & rk_{A \rightarrow B}^{(3)}) \\ &= (H(m||0)^r, & g^r, & H(g^r||1)^{ar'}, & (pk_A)^{r'}, & H((pk_A)^{r'}||2)^{1/b}) \\ &= (H(m||0)^{r_1}, & g^{r_1}, & H(g^{r_1}||1)^{r_2}, & g^{r_2}, & H(g^{r_2}||2)^{1/b}) \end{aligned}$$

Note that we set  $r_1 = r \bmod q$  and  $r_2 = ar' \bmod q$ .

- **Verify:** On input a verification key  $pk$ , a message  $m$  at level  $\ell \in \{1, 2\}$ , and a signature  $\sigma$ ,

- if  $\sigma$  is an *owner-type* signature  $\sigma = (A, B, C)$  (i.e.,  $\ell = 1$ ), it checks

$$e(pk, H(B||1)) \stackrel{?}{=} e(g, C) \text{ and } e(B, H(m||0)) \stackrel{?}{=} e(g, A).$$

If the two equalities both hold, it outputs 1; otherwise, outputs 0.

- if  $\sigma$  is a *nonowner-type* signature  $\sigma = (A, B, C, D, E)$  (i.e.,  $\ell = 2$ ), it

checks

$$\begin{aligned} e(g, H(D||2)) &\stackrel{?}{=} e(pk, E), \\ e(D, H(B||1)) &\stackrel{?}{=} e(g, C), \\ e(B, H(m||0)) &\stackrel{?}{=} e(g, A). \end{aligned}$$

If all the equalities hold, it outputs 1; otherwise, outputs 0.

**Correctness** The correctness is due to the following equalities:

- Owner-type signature:

$$\begin{aligned} e(pk, H(B||1)) &= e(g^a, H(B||1)) = e(g, H(B||1)^a) = e(g, C), \\ e(B, H(m||0)) &= e(g^r, H(m||0)) = e(g, H(m||0)^r) = e(g, A). \end{aligned}$$

- Nonowner-type signature:

$$\begin{aligned} e(pk, E) &= e(g^b, H(D||2)^{1/b}) = e(g, H(D||2)), \\ e(D, H(B||1)) &= e(g^{r_2}, H(B||1)) = e(g, H(B||1)^{r_2}) = e(g, C), \\ e(B, H(m||0)) &= e(g^{r_1}, H(m||0)) = e(g, H(m||0)^{r_1}) = e(g, A). \end{aligned}$$

**Security Analysis of Scheme  $S_{ins}$**  The security proof of scheme  $S_{ins}$  is based on extended computational Diffie–Hellman (eCDH) assumption. That is, it is hard to compute  $g^{uv}$  or  $g^{u/v}$  given  $(q, g, \mathbb{G}, \mathbb{G}_T, e, g^u, g^v, g^{1/v})$ , where  $g$  is a random element in  $\mathbb{G}$ ,  $\mathbb{G}, \mathbb{G}_T$  are bilinear groups with a prime order  $q$ ,  $e$  is a bilinear map,  $\mathbb{G} \times \mathbb{G} \rightarrow \mathbb{G}_T$ , and  $u, v$  are random numbers in  $\mathbb{Z}_q^*$ .

**Theorem 2.2.5.** *Scheme  $S_{ins}$  is secure in the AH model if the eCDH problem is hard, and hash function  $H$  is treated as a random oracle.*

*Proof.* We show that if adversary  $\mathcal{A}$  can break scheme  $S_{ins}$  in the AH model, we can build another algorithm  $\mathcal{B}$  that can solve the eCDH problem. Given

$$(q, g, \mathbb{G}, \mathbb{G}_T, e, g^u, g^v, g^{1/v}),$$

$\mathcal{B}$  aims to output  $g^{uv}$  or  $g^{u/v}$ . The PRS security game goes as follows:

**External Security:**

- *Random oracle  $\mathcal{O}_h$* : On input string  $R$ ,  $\mathcal{B}$  first checks whether  $(R, R_h, r_h, *)$  is in Table  $T_h$ . If yes,  $\mathcal{B}$  returns  $R_h$  and terminates; otherwise,  $\mathcal{B}$  chooses a random number  $r_h \in \mathbb{Z}_q^*$ , and proceeds as follows:
  - The input string  $R$  satisfies the format  $m||0$ , where  $m \in \mathcal{M}$ .  $\mathcal{B}$  guesses whether  $m$  is the target message  $m^*$ . If yes,  $\mathcal{B}$  outputs  $R_h = (g^u)^{r_h}$ ; otherwise,  $\mathcal{B}$  outputs  $R_h = g^{r_h}$ .
  - The input string  $R$  satisfies the format  $m||1$  or  $m||2$ , where  $m \in \mathbb{G}$ .  $\mathcal{B}$  outputs  $R_h = (g^u)^{r_h}$ .
  - The input string  $R$  does not satisfy any of the above formats.  $\mathcal{B}$  outputs  $R_h = g^{r_h}$ .

Finally,  $\mathcal{B}$  records  $(R, R_h, r_h, \perp)$  in Table  $T_h$ .

- *$\mathcal{O}_{pk}$* : On input index  $i$ ,  $\mathcal{B}$  first chooses a random number  $x_i \in \mathbb{Z}_q^*$ , and guesses whether it is  $pk^*$ . For  $pk_i \neq pk^*$ , it sets  $pk_i = g^{x_i}$ ; for  $pk_i = pk^*$ , it sets  $pk_i = (g^v)^{x_i}$ . Finally,  $\mathcal{B}$  records  $(pk_i, x_i)$  in Table  $T_k$ .
- *$\mathcal{O}_s$* : On input  $(pk_i, m_i)$ ,  $\mathcal{B}$  proceeds as follows:
  - If  $m_i = m^*$ , then  $pk_i \neq pk^*$ ,  $\mathcal{B}$  chooses a random number  $r \in \mathbb{Z}_q^*$ , and outputs

$$\sigma = (A, B, C) = (H(m^*||0)^r, g^r, H(g^r||1)^{x_i}).$$

- If  $m_i \neq m^*$ , then  $\mathcal{B}$  chooses a random number  $r \in \mathbb{Z}_q^*$ , and checks whether  $((g^v)^r||1, *, *, *)$  is in Table  $T_h$ . If it exists,  $\mathcal{B}$  outputs **failure** and aborts; otherwise  $\mathcal{B}$  chooses a random number  $r_1 \in \mathbb{Z}_q^*$ , and records  $((g^v)^r||1, g^{r_1}, r_1, r)$  in Table  $T_h$ . Then  $\mathcal{B}$  checks whether  $(m_i||0, \star_1, \star_2, \perp)$

is in Table  $T_h$ . If it exists, then  $\mathcal{B}$  sets  $r_2 = \star_2$ ; otherwise,  $\mathcal{B}$  chooses a random number  $r_2 \in \mathbb{Z}_q^*$  and records  $(m_i || 0, g^{r_2}, r_2, \perp)$  in Table  $T_h$ . Finally,  $\mathcal{B}$  outputs

$$\sigma = (A, B, C) = ((g^v)^{rr_2}, (g^v)^r, pk_i^{r_1}) = (H(m_i || 0)^{vr}, g^{vr}, H(g^{vr} || 1)^{vx_i}).$$

- $\mathcal{O}_{rs}$ : On input  $(pk_i, pk_j, m, \sigma, 1)$ , where  $\sigma = (A, B, C)$ . If  $\text{Verify}(pk_i, m, \sigma, 1) = 1$ , then  $\mathcal{B}$  proceeds as follows; otherwise, outputs **reject**.
  - If  $pk_j \neq pk^*$ ,  $\mathcal{B}$  uses  $x_j$ , associated to  $pk_j$  in Table  $T_k$ , to run **ReKeyGen** and **ReSign**, and gets the required re-signature.
  - If  $pk_j = pk^*$ , then  $m \neq m^*$ , and  $\mathcal{B}$  chooses a random number  $r \in \mathbb{Z}_q^*$ , and checks whether  $((g^v)^{x_i r} || 2, *, *, *)$  is in Table  $T_h$ . If it exists,  $\mathcal{B}$  outputs **failure** and aborts;  $\mathcal{B}$  chooses a random number  $r_1 \in \mathbb{Z}_q^*$ , and records  $((g^v)^{x_i r} || 2, g^{r_1}, r_1, x_i r)$  in Table  $T_h$ .  $\mathcal{B}$  searches  $(B || 1, \star_1, \star_2, \star_3)$  in Table  $T_h$ , and outputs

$$\begin{aligned} \sigma' &= (A', B', C', D', E') \\ &= (A, B, (g^v)^{x_i \star_2 r}, (g^v)^{x_i r}, (g^{1/v})^{r_1/x_t}) \\ &= (A, B, H(B || 1)^{x_i vr}, g^{x_i vr}, H(g^{x_i vr} || 2)^{1/(vx_t)}), \end{aligned}$$

- *Forgery*: At some point, the adversary must output a forgery  $(pk^*, m^*, \sigma^*)$ .

Now, we show how  $\mathcal{B}$  gets the eCDH solution from the forgery as follows.

- If  $\sigma^*$  is an owner-type signature, such that  $\sigma^* = (A^*, B^*, C^*)$ , then we have the following analysis:
  - If any owner-type signature corresponding to  $pk^*$  from  $\mathcal{O}_s$  is not of the form  $(*, B^*, C^*)$ , then  $\mathcal{B}$  finds  $(B^* || 1, \star'_1, \star'_2, \star'_3)$  in Table  $T_h$ , and gets the

solution of the eCDH problem:

$$(C^*)^{1/(x^* \star_2')} = (H(B^* || 1)^{vx_t})^{1/(x^* \star_2')} = (g^{u \star_2' vx_t})^{1/(x^* \star_2')} = g^{uv}.$$

Note that  $pk^* = (g^v)^{x^*}$ .

- Once an owner-type signature corresponding to  $pk^*$  from  $\mathcal{O}_s$  is of the form  $(*, B^*, C^*)$ , then  $\mathcal{B}$  finds out  $(m^* || 0, \star_1, \star_2, \star_3)$  and  $(B^* || 1, \star_1', \star_2', \star_3')$  in Table  $T_h$ , and gets the solution of the eCDH problem:

$$(A^*)^{1/(\star_2 \star_3')} = (H(m^* || 0)^{v \star_3'})^{1/(\star_2 \star_3')} = (g^{uv \star_2 \star_3'})^{1/(\star_2 \star_3')} = g^{uv}.$$

Note that  $B^* = (g^v)^{\star_3'}$ .

- If  $\sigma^*$  is a nonowner-type signature, such that  $\sigma^* = (A^*, B^*, C^*, D^*, E^*)$ , then we have the following analysis:

- If any signature corresponding to  $pk^*$  from  $\mathcal{O}_{rs}$  is not of the special form  $(*, *, *, D^*, E^*)$ ,  $\mathcal{B}$  finds  $(D^* || 2, \star_1'', \star_2'', \star_3'')$  in Table  $T_h$  and gets the solution of the eCDH problem:

$$(E^*)^{x_t / \star_2''} = (H(D^* || 2)^{1/(vx_t)})^{x_t / \star_2''} = (g^{u \star_2'' / (vx_t)})^{x_t / \star_2''} = g^{u/v}.$$

Note that  $pk^* = (g^v)^{x_t}$ .

- If at least one signature corresponding to  $pk^*$  from the output of oracles is of the form  $(*, *, *, D^*, E^*)$ , but any signature corresponding to  $pk^*$  from  $\mathcal{O}_{rs}$  is not of the form  $(*, B^*, C^*, D^*, E^*)$ ,  $\mathcal{B}$  finds out  $(D^* || 2, \star_1'', \star_2'', \star_3'')$  and  $(B^* || 1, \star_1', \star_2', \star_3')$  in Table  $T_h$  and gets the solution of the eCDH problem:

$$(C^*)^{1/(\star_3'' \star_2')} = (H(B^* || 1)^{v \star_3''})^{1/(\star_3'' \star_2')} = (g^{u \star_2' v \star_3''})^{1/(\star_3'' \star_2')} = g^{uv}.$$



Note that  $D^* = (g^v)^{\star_3''}$ .

- Once an owner-type signature corresponding to  $pk^*$  from  $\mathcal{O}_{rs}$  is of the form  $(*, B^*, C^*, D^*, E^*)$ , then  $\mathcal{B}$  finds out  $(m^* || 0, \star_1, \star_2, \star_3)$  and  $(B || 1, \star_1', \star_2', \star_3')$  in Table  $T_h$ , and gets the solution of the eCDH problem:

$$(A^*)^{1/(\star_2 \star_3')} = (H(m^* || 0)^{v \star_3'})^{1/(\star_2 \star_3')} = (g^{uv \star_2 \star_3'})^{1/(\star_2 \star_3')} = g^{uv}.$$

Note that  $B^* = (g^v)^{\star_3'}$ .

Notice that  $\mathcal{B}$  guesses the right target verification key with the probability  $1/q_{pk}$  at least, and  $\mathcal{B}$  outputs **failure** in  $\mathcal{O}_s$  and  $\mathcal{O}_{rs}$  with the probabilities  $(q_h + q_s)/q$  and  $(q_h + q_{rs})/q$  at most, respectively. Here,  $q_{pk}$ ,  $q_h$ ,  $q_s$ , and  $q_{rs}$  are the maximum numbers that  $\mathcal{A}$  can query to  $\mathcal{O}_{pk}$ ,  $\mathcal{O}_h$ ,  $\mathcal{O}_s$ ,  $\mathcal{O}_{rs}$ , respectively. As a result,  $\mathcal{B}$  solves the eCDH problem with a non-negligible probability.

**Internal Security:** Internal security includes limited proxy, delegatee security, and delegator security.

*Limited Proxy:*

- $\mathcal{O}_h$ : Identical to that in external security.
- $\mathcal{O}_{pk}$ : On input index  $i$ ,  $\mathcal{B}$  first chooses a random number  $x_i \in \mathbb{Z}_q^*$ , and then outputs  $pk_i = (g^v)^{x_i}$ . Finally,  $\mathcal{B}$  records  $(pk_i, x_i)$  in Table  $T_k$ .
- $\mathcal{O}_s$ : On input  $(pk_i, m_i)$ ,  $\mathcal{B}$  proceeds as follows:
  - If  $m = m^*$  and  $pk_i \neq pk^*$   $\mathcal{B}$  chooses a random number  $r$ , and checks whether  $(g^r || 1, *, *, *)$  is in Table  $T_h$ . If it exists,  $\mathcal{B}$  outputs **failure** and aborts; otherwise,  $\mathcal{B}$  chooses a random number  $r_1$  and records

$$(g^r || 1, g^{r_1}, r_1, \perp)$$

in Table  $T_h$ . Finally,  $\mathcal{B}$  outputs  $(H(m^*||0)^r, g^r, pk_i^{r_1})$ .

– If  $m \neq m^*$ , then  $\mathcal{B}$  performs the same as that in external security.

- $\mathcal{O}_{rk}$ : On input  $(pk_i, pk_j)$ ,  $\mathcal{B}$  chooses a random number  $r \in \mathbb{Z}_q^*$ , and checks whether  $((g^v)^{x_i r} || 2, *, *, *)$  is in Table  $T_h$ . If it exists,  $\mathcal{B}$  outputs `failure` and aborts; otherwise,  $\mathcal{B}$  chooses a random number  $r_1 \in \mathbb{Z}_q^*$ , records

$$((g^v)^{x_i r} || 2, g^{r_1}, r_1, x_i r)$$

in Table  $T_h$  and outputs  $(r, (g^v)^{x_i r}, (g^{1/v})^{r_1/x_j})$ .

Note that  $pk_i = (g^v)^{x_i}$  and  $pk_j = (g^v)^{x_j}$ .

With the similar analysis in the external security,  $\mathcal{B}$  solves the eCDH problem with a non-negligible probability.

*Delegatee Security:* Compared to the limited proxy,  $\mathcal{B}$  needs to change  $\mathcal{O}_{pk}$ ,  $\mathcal{O}_{sk}$ ,  $\mathcal{O}_s$ , and  $\mathcal{O}_{rk}$  as follows.

- $\mathcal{O}_{pk}$ : For the target delegatee, set the verification key as  $(g^v)^{x_0}$ , and for all other users,  $g^{x_i}$ , where  $x_i (i = 0, \dots, n) \in \mathbb{Z}_q^*$ .
- $\mathcal{O}_{sk}$ : On input  $pk_i$ ,  $\mathcal{B}$  returns the corresponding  $x_i$ .
- $\mathcal{O}_s$ : On input  $(pk_0, m_i)$ ,  $\mathcal{B}$  performs the same as in  $\mathcal{O}_s$  with input  $(pk^*, m_i)$  in external security, where  $pk_0$  is treated as  $pk^*$ .
- $\mathcal{O}_{rk}$ : On input  $(pk_i, pk_j)$ , where  $pk_j \neq pk_0$ ,  $\mathcal{B}$  performs the same as in the real execution since it knows  $x_j$  corresponding to  $pk_j$ .

With the similar analysis in the external security,  $\mathcal{B}$  solves the eCDH problem with a non-negligible probability.

*Delegator Security:* Compared to the limited proxy,  $\mathcal{B}$  needs to change  $\mathcal{O}_{pk}$ ,  $\mathcal{O}_s$ , and  $\mathcal{O}_{rk}$  as follows:

- $\mathcal{O}_{pk}$ : For the target delegator, set the verification key as  $(g^v)^{x_0}$ , and for all other users,  $g^{x_i}$ , where  $x_i (i = 0, \dots, n) \in \mathbb{Z}_q^*$ .
- $\mathcal{O}_{sk}$ : On input  $pk_i, \mathcal{B}$  returns the corresponding  $x_i$ .
- $\mathcal{O}_s$ : On input  $(pk_0, m_i)$ ,  $\mathcal{B}$  performs the same as in  $\mathcal{O}_s$  with input  $(pk^*, m_i)$  in external security, where  $pk_0$  is treated as  $pk^*$ .
- $\mathcal{O}_{rk}$ : On input  $(pk_i, pk_j)$ ,  $\mathcal{B}$  proceeds as follows:
  - If  $pk_j \neq pk_0$ ,  $\mathcal{B}$  gets  $x_j$  from Table  $T_k$ , and uses  $x_j$  to run  $\text{ReKeyGen}(pk_i, pk_j)$ . In the end,  $\mathcal{B}$  outputs the result of  $\text{ReKeyGen}$ .
  - If  $pk_j = pk_0$ ,  $\mathcal{B}$  chooses a random number  $r \in \mathbb{Z}_q^*$  and checks whether  $(pk_i^r || 2, *, *, *)$  is in Table  $T_h$ . If it exists,  $\mathcal{B}$  outputs failure and aborts; otherwise,  $\mathcal{B}$  chooses a random number  $r_1 \in \mathbb{Z}_q^*$ , records  $(pk_i^r || 2, g^{r_1}, r_1, \perp)$  in Table  $T_h$ , and outputs  $(r, pk_i^r, (g^{1/v})^{r_1/x_0})$ .

With the similar analysis in the external security,  $\mathcal{B}$  solves the eCDH problem with a non-negligible probability. Note that in this case, the forgery  $\sigma^*$  is an owner-type signature.

This completes the proof. □

**An Attack on Scheme  $S_{ins}$ :** Although the security proof of scheme  $S_{ins}$  in the AH model is given before, there exists the following attack. Let us now consider the case: Alice  $\rightarrow$  Proxy  $\rightarrow$  Bob, i.e., Bob delegates his signing rights to Alice via Proxy. First, Alice can produce an owner-type signature on  $m$ :  $\sigma_a = (H(m||0)^r, g^r, H(g^r||1)^a)$ , where  $r$  is known to Alice. Then Proxy can transform  $\sigma_a$  into Bob's signature  $\sigma_b = (H(m||0)^r, g^r, (H(g^r||1)^a)^{rk_{a \rightarrow b}^{(1)}}, rk_{a \rightarrow b}^{(2)}, rk_{a \rightarrow b}^{(3)})$ . In this case, Alice can generate signatures on any message simply by changing  $m$  to  $m'$ , since she knows  $r$ . This shows that scheme  $S_{ins}$  is insecure. Hence, the AH model is not suitable for private proxy and unidirectional scheme. Since most of the existing unidirectional PRS

schemes are private proxy and unidirectional schemes, it is desired to propose a new security model to solve this problem.

*Remark 2.2.6.* In fact, the AH model does not consider the attack that the delegatee tries to forge a valid nonowner-type signature of the delegator without colluding with the proxy.

The main reason why scheme  $S_{ins}$  is insecure is that algorithm ReSign does not affect the value  $A$  containing the message  $m$  in the owner-type signature. However, the schemes in [11, 193] do not have this flaw.

### 2.2.6 AH<sup>+</sup> Model

In this section, we propose a new security model for private proxy and unidirectional PRS. Due to its simplicity, it is easy to verify its completeness.

Before giving our security model, we will first define several terms.

1. If user  $A$  delegates his signing rights to user  $B$  via proxy  $P$ , then both users  $A$  and  $B$  are said to be in a *delegation chain*, denoted as  $(B, A)$ . User  $B$  is called user  $A$ 's *delegation predecessor*. The combination of the proxy and a user, either the delegatee  $B$  or the delegator  $A$ , is called a *delegation pair*. Therefore, user  $A$  and proxy  $P$  is a *delegation pair*, so is user  $B$  and proxy  $P$ .
2. If one of the parties in a delegation pair is corrupted, the delegation pair is *corrupted*; otherwise, it is *uncorrupted*.
3. A user can be treated as the smallest *delegation chain*.
4. If two users  $A$  and  $B$  are in a delegation chain and  $B$  is  $A$ 's delegation predecessor,  $B$ 's signature can be transformed by a proxy or proxies into  $A$ 's signature.
5. A delegation chain is its own *subchain*.

6. (Only for multiuse.) If user  $A$  delegates his signing rights to user  $B$  via a proxy  $P$ , and user  $B$  delegates his signing rights to user  $C$  via a proxy  $P'$ , users  $A$  and  $C$  are said to be in a *delegation chain* too. User  $C$  is also called user  $A$ 's *delegation predecessor*. In this case, users  $A, B$ , and  $C$  are in a delegation chain  $(C, B, A)$ . The delegation chains  $(B, A)$  and  $(C, B)$  are *delegation subchains* of the delegation chain  $(C, B, A)$ . The delegation chain  $(C, B, A)$  can be extended if  $C$  delegates his signing rights to another user via another proxy.

### Existential Unforgeability of Private Proxy and Unidirectional Proxy Re-Signature:

The existential unforgeability of private proxy and unidirectional PRS is defined by the following adaptively chosen-message attack game played between a challenger  $\mathcal{C}$  and an adversary  $\mathcal{A}$ . We consider the security in a static situation, that is, before the game starts, the adversary should decide which users and proxies are to be corrupted. Furthermore, the challenger generates all the verification keys in the following game.

**Queries:** The adversary adaptively makes a number of different queries to the challenger. Each query can be one of the following:

- $\mathcal{O}_{pk}$ : On input an index  $i$ , the challenger first runs  $\text{KeyGen}(1^\lambda)$  to get a key pair  $(pk_i, sk_i)$ , and then forwards the verification key  $pk_i$  to the adversary. Finally, the challenger records  $(pk_i, sk_i)$  in Table  $T_k$ . In the following oracles,  $sk_i$  is the signing key corresponding to  $pk_i$ .
- $\mathcal{O}_{sk}$ : On input a verification key  $pk_i$ , the challenger responds with  $sk_i$  which is the associated value with  $pk_i$  in Table  $T_k$ , if  $pk_i$  is corrupted; otherwise, the challenger responds with **reject**.
- $\mathcal{O}_{rk}$ : On input two verification keys  $pk_i, pk_j$  ( $pk_i \neq pk_j$ ), the challenger responds with  $\text{ReKeyGen}(pk_i, pk_j, sk_j)$ .
- $\mathcal{O}_s$ : On input a verification key  $pk_i$  and a message  $m_i$ , the challenger responds with  $\text{Sign}(sk_i, m, 1)$ .

- $\mathcal{O}_{rs}$ : On input two verification keys  $pk_i, pk_j$  ( $pk_i \neq pk_j$ ), a message  $m_i$ , and a signature  $\sigma_i$  at level  $\ell$ , the challenger responds with

$$\text{ReSign}(\text{ReKeyGen}(pk_i, pk_j, sk_j), pk_i, m_i, \sigma_i, \ell).$$

**Forgery:** The adversary outputs a message  $m^*$ , a verification key  $pk^*$ , and a signature  $\sigma^*$  at level  $\ell^*$ . The adversary *wins* if all the following requirements are satisfied:

- $\text{Verify}(pk^*, m^*, \sigma^*, \ell^*) = 1$ .
- $pk^*$  is uncorrupted.
- The adversary has not made a signature query on  $(pk^*, m^*)$ .
- The adversary has not made a signature query on  $(pk', m^*)$ , where  $pk'$  is uncorrupted, and there exists such a delegation subchain from  $pk'$  to  $pk^*$  that does not contain any uncorrupted delegation pair;
- The adversary has not made a re-signature key query on  $(pk_i, pk_j)$ , which satisfies all the following conditions:
  - $pk_i$  is corrupted,
  - $pk_j$  is uncorrupted,
  - There exists such a delegation subchain from  $pk_j$  to  $pk^*$  that does not contain any uncorrupted delegation pair.
- The adversary has not made a re-signature query on  $(pk_i, pk_j, m^*, \sigma_i, *)$ , where  $pk_j$  is uncorrupted, and there exists such a delegation subchain from  $pk_j$  to  $pk^*$  that does not contain any uncorrupted delegation pair.

We define  $\text{Adv}_{\text{PRS}}^{\text{EU}}(\lambda)$  to be the probability that adversary  $\mathcal{A}$  wins in the above game for the security parameter  $\lambda$ .

**Definition 2.2.7.** A private proxy and unidirectional scheme is existentially unforgeable with respect to adaptive chosen message attacks if for all p.p.t. adversaries  $\mathcal{A}$ ,

$\text{Adv}_{\text{PRS}}^{\text{EU}}(\lambda)$  is negligible for the security parameter  $\lambda$ .

*Remark 2.2.8 (Requirement).* The requirements for the adversary's success guarantee that  $(pk^*, m^*, \sigma^*, \ell^*)$  is a valid signature; the adversary cannot trivially obtain a valid owner-type signature by obtaining the signing key; the adversary cannot trivially obtain a valid owner-type signature by the signature oracle; and the adversary cannot trivially obtain a valid nonowner-type signature by the re-signature key oracle and re-signature oracle.

*Remark 2.2.9 (Chosen-Key Model).* Following the spirit in [193], we can easily extend our security model into the chosen-key model [23], where the certificate authority does not need to verify that the owner of one verification key indeed knows the corresponding signing key. In particular, the challenger is no longer responsible for generating the key pairs for corrupted users, and the adversary generates all the corrupted verification keys involved in other oracles.

The static mode can be modified to adaptive mode by removing the restriction that the adversary should decide which users and proxies are to be corrupted before the game starts.

## 2.3 Proxy Re-Encryption

Proxy re-encryption (PRE) is also proposed by Blaze et al. [29] at Eurocrypt 1998. In such an encryption scheme, a semi-trusted proxy, holding some information (a.k.a. re-encryption key), can transform a ciphertext under one public key (delegator Alice) to another ciphertext under another public key (delegatee Bob). These two ciphertexts are corresponding to the same plaintext. However, the proxy cannot obtain the plaintext.

### 2.3.1 Properties and Definition

PRE has the similar desired properties with PRS.

- *Unidirectional*: In this scheme, a re-encryption key allows the proxy to transform Alice's ciphertext to Bob's but not vice versa. In a bidirectional scheme, on the other hand, the re-encryption key allows the proxy to transform Alice's ciphertext to Bob's as well as B's ciphertext to Alice's.
- *Multiuse*: In a multiuse scheme, a transformed ciphertext can be re-transformed again by a proxy. In a single-use scheme, the proxy can transform only the ciphertexts that have not been transformed.
- *Private proxy*: The proxy can keep the re-encryption key as a secret in a private proxy scheme, but anyone can recompute the re-encryption key by observing the process of re-encryption passively in a public proxy scheme.
- *Transparent*: In a transparent scheme, users may not even know the existence of a proxy.
- *Key-optimal*: In a key-optimal scheme, a user is required to protect and store only a small constant amount of secrets no matter how many decryption delegations the user gives or accepts.
- *Noninteractive*: The delegatee is not required to participate in delegation process.
- *Non-transitive*: A re-encryption key cannot be generated from two other re-encryption keys. For example, the re-encryption key from Alice to Bob cannot be generated from the re-encryption keys from Alice to Tina and Tina to Bob.
- *Temporary*: A re-encryption right is temporary. This can be done by either revoking the right [11] or expiring the right.
- *Collusion-resistant*: The delegator can delegate the decryption rights to the delegatee via the proxy, while keeping signing rights for the same public key.



**Definition 2.3.1** (Proxy Re-encryption). A PRE scheme is a tuple of p.p.t. algorithms  $\text{KeyGen}, \text{ReKeyGen}, \text{Enc}, \text{ReEnc}, \text{Dec}$ .

- **KeyGen**: It takes as input the security parameter  $\lambda$ , and returns a public key  $pk$  and a private key  $sk$ .
- **ReKeyGen**: It takes as input delegator Alice's key pair  $(pk_A, sk_A)$  and delegatee Bob's key pair  $(pk_B, sk_B)$ , and returns a re-encryption key  $rk_{A \rightarrow B}$  for the proxy. If the PRE scheme is unidirectional, the delegatee's private key is not included in the input. If the PRE scheme is bidirectional, then the proxy can easily obtain  $rk_{B \rightarrow A}$  from  $rk_{A \rightarrow B}$ . In many bidirectional schemes [65],  $rk_{A \rightarrow B} = 1/rk_{B \rightarrow A}$ .
- **Enc**: It takes as input a public key  $pk$ , a positive integer  $\ell$ , and a message  $m$  from the message space, and returns a ciphertext at level  $\ell$ . If the PRE scheme is single-use, then  $\ell \in \{1, 2\}$ .
- **ReEnc**: It takes as a re-encryption key  $rk_{A \rightarrow B}$  and a ciphertext  $C_A$  under  $pk_A$  at level  $\ell$ , and returns the ciphertext  $C_B$  under  $pk_B$  at level  $\ell + 1$ . If the PRE scheme is single-use, then  $\ell = 1$  and it can be omitted.
- **Dec**: It takes as input a private key  $sk$  and a ciphertext  $C$  at level  $\ell$ , and returns  $m$  in the message space or a special symbol `reject`.

**Correctness.** The correctness property has two requirements. For any message  $m$  in the message space and any key pairs  $(pk, sk), (pk', sk') \leftarrow \text{KeyGen}(1^\lambda)$ . Then the following two conditions must hold:

$$\text{Dec}(sk, \text{Enc}(pk, m, \ell), \ell) = m \text{ and}$$

$$\text{Dec}(sk', \text{ReEnc}(\text{ReKeyGen}(sk, pk'), C, \ell), \ell + 1) = m,$$

where  $C$  is the ciphertext for message  $m$  under  $pk$  from algorithm **Enc** or **ReEnc**.

### 2.3.2 Related Work

Since the introduction of PRE by Blaze *et al.* [29], there have been many papers [7, 9, 10, 29, 65, 86, 96, 130, 145, 156, 246] that have proposed different PRE schemes with different security properties.

The first bidirectional, multiuse PRE scheme secure against chosen-plaintext attack (CPA) was proposed by Blaze *et al.* [29]; however, it is not collusion resistant. Based on public key encryption with double trapdoors, Ateniese *et al.* [9, 10] proposed the first collusion resistant, unidirectional, single-use PRE schemes with CPA security. At TCC 2007, Hohenberger *et al.* [145] proposed a new collusion resistant, CPA-secure, unidirectional PRE scheme, where the re-encryption key together with the re-encryption algorithm can be treated as an obfuscated re-encryption program.

However, many PRE's applications, such as the distributed file system, demand that the underlying PRE scheme is secure against chosen-ciphertext attack (CCA). To solve the problem, Canetti and Hohenberger [65] and Ateniese and Green [130] proposed the definition of CCA security independently. There are two kinds of CCA-secure PRE schemes: with pairings and without pairings.

By using pairings and the CHK paradigm [36, 63], Canetti and Hohenberger [65] proposed the first CCA-secure (bidirectional) PRE scheme in the standard model. However, their scheme suffers from collusion attacks. Furthermore, they didn't propose any CCA-secure unidirectional PRE scheme, and left it as an open problem. Based on Canetti-Hohenberger technique, Libert and Vergnaud [194, 195] proposed a new unidirectional PRE scheme that is CCA-secure and collusion resistant in the standard model. Recently, Weng *et al.* [293] and we [248] proposed CCA-secure and collusion-resistant unidirectional PRE schemes by improving Libert and Vergnaud's method.

Due to the heavy cost of pairing computation, it is desired to design CCA-secure PRE schemes without pairings. We [246] and Chow *et al.* [86] proposed CCA-secure and collusion-resistant unidirectional PRE schemes without pairings. However, the pro-

posed schemes are only proven secure in the random oracle model. Most recently, Matsuda *et al.* [210] proposed a new pairing-free, CCA-secure (bidirectional) PRE scheme, which is proven secure in the standard model. However, it suffers from the collusion attack. Furthermore, Weng *et al.* [295] pointed out that the scheme of Matsuda *et al.* is not CCA-secure, but they did not suggest any improvement.

In 2009<sup>1</sup>, we [249] proposed a generic construction for CCA-secure, single-use, unidirectional PRE from CCA-secure (2,2) threshold public key encryption. With the generic construction, we can obtain two kinds of single-use, unidirectional PRE schemes. One is secure in the random oracle model without pairings, and the other is secure in the standard model with pairings. At CT-RSA 2012, Hanaoka *et al.* [136] obtained a similar result by using similar methods.

There are also some PRE schemes with some special properties, such as threshold [249], conditional [255, 272, 294], invisible proxy [160, 272], identity-based [130, 255, 283], anonymous [7, 253, 254], and searchable [247].

In the rest of this section, we first introduce the security models of (unidirectional) PRE. In what follows, we introduce the PRE scheme [246] proposed at PKC 2009.

### 2.3.3 Security Models

Usually, unidirectional PRE is better than bidirectional PRE. For example, the delegator delegates his decryption rights to the delegatee, the delegatee does not always want to do the reverse delegation. Furthermore, any unidirectional scheme can be easily transformed to a bidirectional one by running the former in both directions, while whether the reverse holds is unknown. Hence, we only give the security models for single-use unidirectional PRE and multiuse unidirectional PRE in this section.

The security models below are defined by games played between a challenger  $\mathcal{C}$  and an adversary  $\mathcal{A}$ . We assume that the public keys input into the oracles by the adversary

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<sup>1</sup>The paper [249] was published online in May, 2009.

are all from public key generation oracle  $\mathcal{O}_{pk}$ . We say that a public key is corrupted if it has been queried to private key generation oracle  $\mathcal{O}_{sk}$  by the adversary.

### Chosen Ciphertext Security for Single-use, Unidirectional Proxy Re-encryption (SUPRE)

In almost all existing unidirectional PRE schemes except one [119], there are two types of ciphertexts. One is original ciphertext ( $\ell = 1$ ) that can only be generated by  $\text{Enc}$ , and the other is re-encrypted ciphertext ( $\ell = 2$ ) that can be generated by not only  $\text{Enc}$  but also  $\text{ReEnc}$ . Hence, we have two security models for SUPRE, one aims to protect the plaintext corresponding to the original ciphertext, and the other aims to protect the plaintext corresponding to the re-encrypted ciphertext.

#### The challenge ciphertext is an original ciphertext.

**Setup:** The challenger sets up the system parameters.

**Phase 1:** The adversary  $\mathcal{A}$  can issue the following queries adaptively.

- *Public key generation oracle  $\mathcal{O}_{pk}$ :* The challenger takes a security parameter  $k$ , runs  $\text{KeyGen}(1^\lambda)$  to generate a key pair  $(pk_i, sk_i)$ , gives  $pk_i$  to  $\mathcal{A}$  and records  $(pk_i, sk_i)$  in Table  $T_k$ . In the following,  $sk_i$  is the corresponding private key to  $pk_i$ .
- *Private key generation oracle  $\mathcal{O}_{sk}$ :* On input  $pk_i$ , the challenger searches for  $pk_i$  in Table  $T_k$  and returns  $sk_i$ .
- *Re-encryption key generation oracle  $\mathcal{O}_{rk}$ :* On input  $(pk_i, pk_j)$ , the challenger returns the re-encryption key  $rk_{i,j} = \text{ReKeyGen}(sk_i, pk_j)$ .
- *Re-encryption oracle  $\mathcal{O}_{re}$ :* On input  $(pk_i, pk_j, C)$ , the challenger returns the re-encrypted ciphertext  $C' = \text{ReEnc}(\text{ReKeyGen}(sk_i, pk_j), C)$ .

- *Decryption oracle*  $\mathcal{O}_{dec}$ : On input  $(pk_i, C_i)$ , the challenger returns  $\text{Dec}(sk_i, C_i)$ .

**Challenge:** Once Phase 1 is over,  $\mathcal{A}$  outputs two plaintexts  $m_0, m_1$  with equal length from the message space, and a public key  $pk^*$  on which  $\mathcal{A}$  wishes to challenge. There are two restrictions on the public key  $pk^*$ : (i)  $pk^*$  has not been queried to  $\mathcal{O}_{sk}$ ; (ii) if  $(pk^*, \star)$  has appeared in any query to  $\mathcal{O}_{rk}$ ,  $\star$  should not be queried to  $\mathcal{O}_{sk}$ . The challenger picks a random bit  $b \in \{0, 1\}$  and sets  $C^* = \text{Enc}(pk^*, m_b)$ . It sends  $C^*$  as the challenge to  $\mathcal{A}$ .

**Phase 2:** This phase is almost the same as Phase 1 but with the following constraints:

- $\mathcal{O}_{sk}$ : On input  $pk_i$ , if  $pk_i = pk^*$ , the challenger outputs **reject**.
- $\mathcal{O}_{rk}$ : On input  $(pk_i, pk_j)$ , if  $pk_i = pk^*$ , and  $pk_j$  is corrupted, the challenger outputs **reject**.
- $\mathcal{O}_{re}$ : On input  $(pk_i, pk_j, C_i)$ , if  $(pk_i, C_i) = (pk^*, C^*)$  and  $pk_j$  is corrupted, the challenger outputs **reject**.
- $\mathcal{O}_{dec}$ : On input  $(pk_i, C_i)$ , if  $(pk_i, C_i)$  is a derivative (see the part after Definition 2.3.2 below) of  $(pk^*, C^*)$ , the challenger outputs **reject**.

**Guess:** Finally, the adversary  $\mathcal{A}$  outputs a guess  $b' \in \{0, 1\}$  and wins the game if  $b = b'$ .

We refer to such an adversary  $\mathcal{A}$  as a CCA-O adversary. We define adversary  $\mathcal{A}$ 's advantage in attacking SUPRE as the following function:

$$\text{Adv}_{\text{SUPRE}}^{\text{CCA-O}}(\lambda) = |\Pr[b = b'] - 1/2|.$$

Using the CCA-O game, we can define CCA-O security of SUPRE.

**Definition 2.3.2.** [CCA-O Security] We say that a single-use, unidirectional PRE scheme SUPRE is CCA-O-secure, if for any polynomial time CCA-O adversary  $\mathcal{A}$ , the function  $\text{Adv}_{\text{SUPRE}}^{\text{CCA-O}}(\lambda)$  is negligible.

The core of PRE's security is the definition of derivatives of ciphertext. Derivatives of  $(pk^*, C^*)$  could be obtained by combining some of the following rules:

**R-1**  $(pk^*, C^*)$  is a derivative of itself.

**R-2** If  $\mathcal{A}$  has queried  $\mathcal{O}_{re}$  on input  $(pk^*, pk, C^*)$  and obtained  $C$ ,  $(pk, C)$  is a derivative of  $(pk^*, C^*)$ .

**R-3** If  $\mathcal{A}$  has queried  $\mathcal{O}_{rk}$  on input  $(pk^*, pk)$  and  $C \leftarrow \text{ReEnc}(\mathcal{O}_{rk}(pk^*, pk), C^*)$ ,  $(pk, C)$  is a derivative of  $(pk^*, C^*)$ .

**R-4** If  $\mathcal{A}$  has queried  $\mathcal{O}_{rk}$  on input  $(pk^*, pk)$  and  $\text{Dec}(pk, C) \in \{m_0, m_1\}$ ,  $(pk, C)$  is a derivative of  $(pk^*, C^*)$ .

**R-5** If  $(pk, C)$  is a re-encrypted ciphertext and  $\text{Dec}(pk, C) \in \{m_0, m_1\}$ ,  $(pk, C)$  is a derivative of  $(pk^*, C^*)$ .

The definition of derivatives we use in this section consists of relations R-1, R-2, and R-3, named  $\text{Def}_{\text{der}}$ . The single-use unidirectional case extended from the Canetti-Hohenberger definition [65] corresponds to relations R-1, R-2, and R-4, and the Libert-Vergnaud definition [194, 195] consists of items R-1 and R-5.

R-1 is the trivial reflexive relation. R-2 models the direct consequence of the re-encryption process. R-5 follows the spirit of the definition of re-randomizable CCA security for the re-encrypted ciphertext. One possible stronger rule (as R-4) is that only when  $\mathcal{A}$  has queried the required re-encryption key, then all re-encrypted ciphertexts that are generated by such a key and have either one of the challenge messages as the plaintext are disallowed to be queried to  $\mathcal{O}_{dec}$ . R-3, in a strict sense, is still deviated from the regular CCA security since a class of ciphertexts instead of a single ciphertext is disallowed.

Most importantly,  $\text{Def}_{\text{der}}$  disallows a *smaller* class of ciphertexts when compared with other definitions. Consider an adversary which has queried the re-encryption oracle

with  $(pk^*, pk, C^*)$  and obtained  $(pk, C)$ , where  $pk$  is an uncorrupted user. It is possible that this adversary can change  $C$  into  $C'$  while the adversary does not require to run ReEnc algorithm, yet the output of  $\text{Dec}(C, pk)$  equals to the one of  $\text{Dec}(C', pk)$ , or it can be computed easily from the one of  $\text{Dec}(C', pk)$ . According to the extension of the Canetti-Hohenberger definition and the Libert-Vergnaud definition,  $(pk, C')$  is a derivative of  $(pk^*, C^*)$ . However,  $(pk, C')$  is *not* a derivative of  $(pk^*, C^*)$  in  $\text{Def}_{\text{der}}$ .

Essentially,  $\text{Def}_{\text{der}}$  *only* takes the re-encrypted ciphertext obtained *directly* either from the algorithm ReEnc or the oracle  $\mathcal{O}_{re}$  as the derivative. This feature makes  $\text{Def}_{\text{der}}$  closer to the CCA security definition for the typical public key encryption.

**The challenge ciphertext is a re-encrypted level ciphertext.** The CCA-R security of SUPRE is defined by the same approach of the CCA-O security.

**Phase 1, Guess:** Identical to that in CCA-O game.

**Challenge:** Once Phase 1 is over,  $\mathcal{A}$  outputs two plaintexts  $m_0, m_1$  with equal length from the message space, and two uncorrupted public keys  $pk, pk^*$  on which  $\mathcal{A}$  wishes to challenge. The challenger picks a random bit  $\mathbf{b} \in \{0, 1\}$  and sets  $C^* = \text{ReEnc}(rk, \text{Enc}(pk, m_{\mathbf{b}}))$ , where  $rk$  is a re-encryption key from  $pk$  to  $pk^*$ . It sends  $C^*$  as the challenge to  $\mathcal{A}$ .

**Phase 2:** Almost the same as in Phase 1 but with the following constraints:

- $\mathcal{O}_{sk}$ : On input  $pk_i$ , if  $pk_i = pk$  or  $pk_i = pk^*$ , the challenger outputs reject.
- $\mathcal{O}_{dec}$ : On input  $(pk_i, C_i)$ , if  $(pk_i, C_i) = (pk^*, C^*)$ , the challenger outputs reject.

We also define  $\text{Adv}_{\text{SUPRE}}^{\text{CCA-R}}(\lambda) = |\Pr[\mathbf{b} = \mathbf{b}'] - 1/2|$  for the security parameter  $\lambda$  as in CCA-O security.

**Definition 2.3.3** (CCA-R Security). We say that a single-use, unidirectional PRE scheme  $\text{SUPRE}$  is CCA-R-secure, if for any polynomial time CCA-R adversary  $\mathcal{A}$  the function  $\text{Adv}_{\text{SUPRE}}^{\text{CCA-R}}(\lambda)$  is negligible.

*Remark 2.3.4.* There is another security, named collusion resistance, guaranteeing that any collusion of delegatee and proxy cannot lead to the reveal of the delegator's private key. It is easy to see that the CCA-R security implies collusion resistance, since anyone obtaining the delegator's private key can definitely decrypt the re-encrypted ciphertext for the delegator.

### Chosen-Ciphertext Security for Multiuse, Unidirectional Proxy Re-Encryption

Before giving our security model, we will first define several terms for PRE.

- If user  $A$  delegates his decryption rights to user  $B$  via a proxy  $P$ , then both users  $A$  and  $B$  are said to be in a *delegation chain*, denoted as  $(A, B)$ . If user  $B$  delegates his decryption rights to user  $C$  via a proxy  $P'$ , then users  $A$  and  $C$  are said to be in a *delegation chain* too, the resulting delegation chain is denoted as  $(A, B, C)$ . If user  $C$  delegates his decryption rights to other people, the delegation chain is further extended.
- One user is the smallest delegation chain.
- If the first and last user in a delegation chain are users  $A$  and  $B$ , respectively, then the delegation chain can be denoted as  $(A, \dots, B)$ .
- For a delegation chain  $(A, \dots, B)$ , if the ciphertext for user  $A$  can be transformed to the ciphertext for user  $B$  by the outputs of the re-encryption key generation oracle and the private key generation oracle, the delegation chain  $(A, \dots, B)$  is *corrupted*; otherwise, it is *uncorrupted*.

Now, we give the CCA security for multiuse, unidirectional PRE  $\text{MUPRE}$  by the following CCA game.



**Setup:**  $\mathcal{C}$  sets up the system parameters, and sends the system parameters to  $\mathcal{A}$ .

**Phase 1:**  $\mathcal{A}$  adaptively issues the following queries:

- *Public key generation oracle*  $\mathcal{O}_{pk}$ : On input an index  $i$ ,  $\mathcal{C}$  first takes a security parameter  $1^\lambda$ , and then runs algorithm  $\text{KeyGen}(1^\lambda)$  to generate a key pair  $(pk_i, sk_i)$ , gives  $pk_i$  to  $\mathcal{A}$  and records  $(pk_i, sk_i)$  in Table  $T_k$ . In the following oracles,  $sk_i$  is the private key corresponding to  $pk_i$ .
- *Private key generation oracle*  $\mathcal{O}_{sk}$ : On input  $pk_i$ ,  $\mathcal{C}$  searches  $pk_i$  in Table  $T_k$  and returns the associated  $sk_i$ .
- *Re-encryption key generation oracle*  $\mathcal{O}_{rk}$ : On input  $(pk_i, pk_j)$ ,  $\mathcal{C}$  returns the re-encryption key  $rk_{i,j} = \text{ReKeyGen}(sk_i, pk_j)$ .
- *Re-encryption oracle*  $\mathcal{O}_{re}$ : On input  $(pk_i, pk_j, (\ell, C))$ ,  $\mathcal{C}$  returns the re-encrypted ciphertext

$$(\ell + 1, C') \leftarrow \text{ReEnc}(\text{ReKeyGen}(sk_i, pk_j), (\ell, C)).$$

- *Decryption oracle*  $\mathcal{O}_{dec}$ : On input  $(pk, (\ell, C))$ ,  $\mathcal{C}$  returns  $\text{Dec}(sk, (\ell, C))$ .

**Challenge:** Once Phase 1 is over,  $\mathcal{A}$  outputs two plaintexts  $m_0, m_1$  with equal length from the message space, a positive integer  $\ell^*$ , a set of public keys  $\{pk_{i_j}\}_{j=1}^{\ell^*-1}$ , and an uncorrupted public key  $pk^*$  on which  $\mathcal{A}$  wishes to challenge. There is one constraint on the public key  $pk^*$ , i.e., there does not exist a corrupted delegation chain from  $pk^*$  to a corrupted public key.  $\mathcal{C}$  picks a random bit  $\mathbf{b} \in \{0, 1\}$  and sets

$$C^* = \text{ReEnc}(\text{ReKeyGen}(sk^*, pk_{i_{\ell^*-1}}), \text{ReEnc}(\text{ReKeyGen}(sk_{i_{\ell^*-1}}, pk_{i_{\ell^*-2}}), \dots, \text{ReEnc}(\text{ReKeyGen}(sk_{i_1}, pk_{i_2}), \text{Enc}(pk_1, m_{\mathbf{b}}))))).$$

It sends  $(\ell^*, C^*)$  as the challenge to  $\mathcal{A}$ .

Note that when  $\ell^* = 1$ , then  $\{pk_{i_j}\}_{j=1}^{\ell^*-1}$  is an empty set, and  $C^* = \text{Enc}(pk^*, m_b)$ .

**Phase 2:**  $\mathcal{A}$  adaptively issues more queries:

- $\mathcal{O}_{pk}$ : Identical to that in Phase 1.
- $\mathcal{O}_{sk}$ : On input  $pk_i$ , if there exists a corrupted delegation chain from  $pk^*$  to  $pk_i$ ,  $\mathcal{C}$  outputs `reject`; otherwise,  $\mathcal{C}$  executes the same steps as in Phase 1.
- $\mathcal{O}_{rk}$ : On input  $(pk_i, pk_j)$ , if one of the following conditions holds,  $\mathcal{C}$  outputs `reject`; otherwise,  $\mathcal{C}$  executes the same steps as in Phase 1.
  - $pk_i$  is uncorrupted and  $pk_j$  is corrupted, and there is a corrupted delegation chain from  $pk^*$  to  $pk_i$ .
  - $pk_i$  is uncorrupted and  $pk_j$  is corrupted, and there is a corrupted delegation chain from  $pk$  to  $pk_i$ , where the adversary has a derivative (see below) of  $(pk^*, (\ell^*, C^*))$ :  $(pk, (\ell, C))$ .

The definition of derivatives of  $(pk^*, (\ell^*, C^*))$  is different from that in single-use case.

1.  $(pk^*, (\ell^*, C^*))$  is a derivative of itself.
  2. If  $(pk, (\ell, C))$  is a derivative of  $(pk^*, (\ell^*, C^*))$  and  $(pk', (\ell', C'))$  is a derivative of  $(pk, (\ell, C))$ ,  $(pk', (\ell', C'))$  is a derivative of  $(pk^*, (\ell^*, C^*))$ .
  3. If  $(pk', (\ell', C')) \leftarrow \mathcal{O}_{re}(pk, pk', (\ell, C))$ ,  $(pk', (\ell', C'))$  is a derivative of  $(pk, (\ell, C))$ .
  4. If  $(\ell', C') \leftarrow \text{ReEnc}(\mathcal{O}_{rk}(pk, pk'), (\ell, C))$ ,  $(pk', (\ell', C'))$  is a derivative of  $(pk, (\ell, C))$ .
  5. If  $(\ell', C') \leftarrow \text{ReEnc}(\text{ReKeyGen}(sk, pk'), (\ell, C))$ ,  $(pk', (\ell', C'))$  is a derivative of  $(pk, (\ell, C))$ , where  $sk$  is the private key of  $pk$ .
- $\mathcal{O}_{re}$ : On input  $(pk_i, pk_j, (\ell, C))$ , if  $(pk_i, (\ell, C))$  is a derivative of  $(pk^*, (\ell^*, C^*))$ , and there is a corrupted delegation chain from  $pk_j$  to a

corrupted public key or  $pk_j$  is corrupted,  $\mathcal{C}$  outputs **reject**; otherwise,  $\mathcal{C}$  responds the same as in Phase 1.

- $\mathcal{O}_{dec}$ : On input  $(pk_i, (\ell, C))$ , if  $(pk_i, (\ell, C))$  is a derivative of  $(pk^*, (\ell^*, C^*))$ ,  $\mathcal{C}$  outputs **reject**; otherwise,  $\mathcal{C}$  responds the same as in Phase 1.

**Guess:** Finally,  $\mathcal{A}$  outputs a guess  $\mathbf{b}' \in \{0, 1\}$  and wins the game if  $\mathbf{b} = \mathbf{b}'$ .

We also define  $\text{Adv}_{\text{MUPRE}}^{\text{CCA}}(\lambda) = |\Pr[\mathbf{b} = \mathbf{b}'] - 1/2|$  for the security parameter  $\lambda$  as that in CCA-O security for SUPRE.

**Definition 2.3.5** (CCA Security). We say that a multiuse, unidirectional PRE scheme MUPRE is CCA-secure, if for any polynomial time CCA adversary  $\mathcal{A}$  the function  $\text{Adv}_{\text{MUPRE}}^{\text{CCA}}(\lambda)$  is negligible.

### Collusion Resistance for Multiuse, Unidirectional Proxy Re-encryption

We still use the above method to define collusion resistance. In the CR game, there exist the same oracles as that in Phase 1 of the CCA game for multiuse, unidirectional PRE. In the end, the adversary wins the game if it outputs the private key of an uncorrupted user. The adversary is named as a CR adversary.

**Definition 2.3.6.** We say that a multiuse, unidirectional PRE scheme MUPRE is collusion resistant, if for any polynomial time CR adversary  $\mathcal{A}$  the function  $\text{Adv}_{\text{MUPRE}}^{\text{CCA-O}}(\lambda)$  is negligible.

*Remark 2.3.7* (CCA Security vs. Collusion Resistance). Unlike the single-use unidirectional proxy, the CCA security (for  $\ell$ -level ciphertext,  $\ell > 1$ ) of multiuse, unidirectional PRE cannot guarantee the CR security, since in the CCA game, the adversary cannot query  $\mathcal{O}_{rk}$  to get the re-encryption key from the target user to any corrupted users, but it is allowed in the CR game.

### 2.3.4 Single-Use, Unidirectional Scheme

From the security model of SUPRE, we see that there are three sufficient conditions for CCA-secure SUPRE:

- 1) Each of the proxy and delegator can verify the validity of the original ciphertext.
- 2) The delegatee can verify the validity of the re-encrypted ciphertext.
- 3) The delegatee colluding with the proxy cannot get the private key of the delegator.

The first condition can be easily met if the original ciphertext is publicly verifiable. There are many methods to obtain public verifiability, such as signature of knowledge. There are as well many methods to get the second condition such as Fujisaki-Okamoto conversion [115]. The last condition is usually the most difficult one to realize on designing CCA-secure unidirectional PRE. From the definition of PRE, the delegatee colluding with the proxy can definitely get a value that can be used to do decryption as the private key of the delegator. Hence, the value, obtained by the collusion of the delegatee and proxy, should not be the real private key of the delegator, instead a sub-private key. The similar concept of sub-private key has appeared in other cryptographic primitives, such as public key encryption with double trapdoors (both trapdoors can be used to do decryption, and one trapdoor can be computed from another one but not vice versa).

Following the above method, we [246] proposed the first CCA-secure, single-use, unidirectional PRE scheme  $S_{\text{SUPRE}}$  in the random oracle model by using Fujisaki-Okamoto conversion, signature of knowledge and the public key encryption with double trapdoors proposed by Bresson et al. [54] (named BCP03).

#### Signature of Knowledge

The following noninteractive zero-knowledge proof of knowledge, named signature of knowledge of equality of two discrete logarithms [8, 61, 258], will be used in followings:

**Definition 2.3.8.** Let  $y_1, y_2, g, h \in \mathbb{G}$ ,  $\mathbb{G}$  be a cyclic group of quadratic residues modulo  $N^2$  ( $N$  is a safe-prime modulus), and  $H(\cdot) : \{0, 1\}^* \rightarrow \{0, 1\}^\lambda$  ( $\lambda$  is the security parameter). A pair  $(c, s)$ , verifying  $c = H(y_1 || y_2 || g || h || g^s y_1^c || h^s y_2^c || m)$  is a signature of knowledge of the discrete logarithm of both  $y_1 = g^x$  w.r.t. base  $g$  and  $y_2 = h^x$  w.r.t. base  $h$ , on a message  $m \in \{0, 1\}^*$ .

The party in possession of the secret  $x$  is able to compute the signature, provided that  $x = \log_g y_1 = \log_h y_2$ , by choosing a random  $t \in \{0, \dots, 2^{|N^2|+k} - 1\}$  ( $|n|$  is the bit-length of  $n$ ). And then computing  $c$  and  $s$  as:

$$c = H(y_1 || y_2 || g || h || g^t || h^t || m) \text{ and } s = t - cx.$$

We denote  $\text{SoK.Gen}(y_1, y_2, g, h, m)$  as the generation of the proof.

### The Public Key Encryption with Double Trapdoors: BCP03

The following description is from [54]. Let  $N = pq$  be a safe prime modulus like  $p = 2p' + 1$ ,  $q = 2q' + 1$ , and  $p, p', q, q'$  are primes. Assume  $\mathbb{G}$  is the cyclic group of quadratic residues modulo  $N^2$ , then the order of  $\mathbb{G}$  is  $Np'q'$ .

- **KeyGen:** Choose a random element  $\alpha \in \mathbb{Z}_{N^2}^*$ , a random value  $a \in [1, Np'q']$ , and set  $g = \alpha^2 \bmod N^2$  and  $h = g^a \bmod N^2$ . The public key is  $(N, g, h)$  and the private key is  $a$ .
- **Enc:** On input a public key  $pk$  and a message  $m \in \mathbb{Z}_N$ , the ciphertext  $(A, B)$  is computed as

$$A = g^r \bmod N^2, \quad B = h^r(1 + mN) \bmod N^2,$$

where  $r$  is a random number from  $\mathbb{Z}_{N^2}$ .

- Dec: There are two methods to decrypt:

- Knowing  $a$ , one can compute  $m$  by

$$m = \frac{B/(A^a) - 1 \bmod N^2}{N}.$$

- Knowing  $p', q'$ , one can compute  $m$  by

$$m = \frac{D - 1 \bmod N^2}{N} \cdot \pi \bmod N,$$

where  $D = \left(\frac{B}{g^{w_1}}\right)^{2p'q'}$ ,  $w_1 = ar \bmod N$ ,  $ar \bmod pqp'q' = w_1 + w_2N$ ,  $\pi$  is the inverse of  $2p'q' \bmod N$ .

Note that the values of  $a \bmod N$  and  $r \bmod N$  can be computed given  $h = g^a \bmod N^2$ ,  $A = g^r \bmod N^2$ , and  $p', q'$ , by the method in [225] (Theorem 1 in [225]).

### The Description of Scheme $S_{\text{SUPRE}}$

Scheme  $S_{\text{SUPRE}}$  contains three cryptographic hash functions:  $H_1(\cdot) : \{0, 1\}^* \rightarrow \{0, 1\}^{\lambda_1}$ ,  $H_2(\cdot) : \{0, 1\}^* \rightarrow \{0, 1\}^n$ , and  $H_3(\cdot) : \{0, 1\}^* \rightarrow \{0, 1\}^{\lambda_2}$ , where  $\lambda_1$  and  $\lambda_2$  are the security parameters,  $n$  is the bit-length of messages to be encrypted. The details are as follows:

- KeyGen: Choose a safe-prime modulus  $N = pq$ , three random numbers  $\alpha \in \mathbb{Z}_{N^2}^*$ ,  $a, b \in [1, pp'qq']$ , a hash function  $H(\cdot)$ , where  $p = 2p' + 1$ ,  $q = 2q' + 1$ ,  $p, p', q, q'$  are primes, and  $H(\cdot) : \{0, 1\}^* \rightarrow \mathbb{Z}_{N^2}$ . Furthermore, set  $g_0 = \alpha^2 \bmod N^2$ ,  $g_1 = g_0^a \bmod N^2$ , and  $g_2 = g_0^b \bmod N^2$ . The public key is  $pk = (H(\cdot), N, g_0, g_1, g_2)$ , the sub-private key is  $wsk = (a, b)$ , and the private key is  $sk = (p, q, p', q')$ .

- **ReKeyGen:** On input a public key  $pk_Y = (H_Y(\cdot), N_Y, g_{Y0}, g_{Y1}, g_{Y2})$ , a sub-private key  $usk_X = a_X$ , and a private key  $sk_X = (p_X, q_X, p'_X, q'_X)$ , it outputs the unidirectional re-encryption key  $rk_{X \rightarrow Y} = (rk_{X \rightarrow Y}^{(1)}, rk_{X \rightarrow Y}^{(2)})$ , where  $rk_{X \rightarrow Y}^{(1)} = (\dot{A}, \dot{B}, \dot{C})$ , and computed as follows:

- Choose two random numbers  $\dot{\sigma} \in \mathbb{Z}_N, \dot{\beta} \in \{0, 1\}^{\lambda_1}$ .
- Compute  $rk_{X \rightarrow Y}^{(2)} = a_X - \dot{\beta} \bmod (p_X q_X p'_X q'_X)$ .
- Compute  $r_{X \rightarrow Y} = H_Y(\dot{\sigma} || \dot{\beta}), \dot{A} = (g_{Y0})^{r_{X \rightarrow Y}} \bmod (N_Y)^2, \dot{C} = H_1(\dot{\sigma}) \oplus \dot{\beta},$

$$\dot{B} = (g_{Y2})^{r_{X \rightarrow Y}} \cdot (1 + \dot{\sigma} N_Y) \bmod (N_Y)^2 \quad (2.1)$$

- **Enc:** On input a public key  $pk = (H(\cdot), N, g_0, g_1, g_2)$  and a message  $m \in \{0, 1\}^n$ , it proceeds as follows:

- Choose a random number  $\sigma \in \mathbb{Z}_N$ .
- Compute  $r = H(\sigma || m), A = (g_0)^r \bmod N^2, C = H_2(\sigma) \oplus m, D = (g_2)^r \bmod N^2,$

$$B = (g_1)^r \cdot (1 + \sigma N) \bmod N^2. \quad (2.2)$$

- Run  $(c, s) \leftarrow \text{SoK.Gen}(A, D, g_0, g_2, (B, C))$ , where the underlying hash function is  $H_3$ .
- Output the ciphertext  $K = (A, B, C, D, c, s)$ .
- **ReEnc:** On input a re-encryption key  $rk_{X \rightarrow Y} = (rk_{X \rightarrow Y}^{(1)}, rk_{X \rightarrow Y}^{(2)})$  and a ciphertext  $K = (A, B, C, D, c, s)$  under key  $pk_X = (H_X(\cdot), N_X, g_{X0}, g_{X1}, g_{X2})$ , check whether  $c = H_3(A || D || g_{X0} || g_{X2} || (g_{X0})^s A^c || (g_{X2})^s D^c || (B || C))$ . If not hold, output **reject** and terminate; otherwise, re-encrypt the ciphertext to be one under key  $pk_Y$  as:

- Compute  $A' = A^{rk_{X \rightarrow Y}^{(2)}} = (g_{X0})^{r(a_X - \dot{\beta})} \bmod (N_X)^2$ .
- Output the new ciphertext  $(A, A', B, C, rk_{X \rightarrow Y}^{(1)}) = (A, A', B, C, \dot{A}, \dot{B}, \dot{C})$ .
- Dec: On input a private key and any ciphertext  $K$ , parse  $K = (A, B, C, D, c, s)$ , or  $K = (A, A', B, C, \dot{A}, \dot{B}, \dot{C})$ .
  - Case  $K = (A, B, C, D, c, s)$ : Check whether

$$c = H_3(A || D || g_0 || g_2 || (g_0)^s A^c || (g_2)^s D^c || (B || C)),$$

if not, output **reject** and terminate; otherwise,

- \* compute  $\sigma = \frac{B/(A^a) - 1 \bmod N^2}{N}$ , if the input private key is the sub-private key  $a$ ;
- \* compute  $\sigma = \frac{(B/g_0^{w_1})^{2p'q'} - 1 \bmod N^2}{N} \cdot \pi(\bmod N)$ , where  $w_1$  is computed as that in scheme BCP03, and  $\pi$  is the inverse of  $2p'q' \bmod N$ , if the private key is the private key  $(p, q, p', q')$ .

Compute  $m = C \oplus H_2(\sigma)$ , if  $A = (g_0)^{H(\sigma || m)}$  and  $B = (g_1)^{H(\sigma || m)} \cdot (1 + \sigma N) \bmod N^2$  both hold, output  $m$ ; otherwise, output **reject** and terminate.

- Case  $K = (A, A', B, C, \dot{A}, \dot{B}, \dot{C})$ : In this case, the decryptor should know the delegator's (Alice's) public key  $(H'(\cdot), N', g'_0, g'_1, g'_2)$ .
  - \* If the input private key is the sub-private key  $b$ , compute

$$\dot{\sigma} = \frac{\dot{B}/(\dot{A}^b) - 1 \bmod N^2}{N}.$$

- \* If the input private key is the private key  $(p, q, p', q')$ , computes

$$\dot{\sigma} = \frac{(\dot{B}/g_0^{w_1})^{2p'q'} - 1 \bmod N^2}{N} \cdot \pi(\bmod N),$$



where  $w_1$  is computed as that in scheme BCP03, and  $\pi$  is the inverse of  $2p'q' \bmod N$ .

Compute  $\dot{\beta} = \dot{C} \oplus H_1(\dot{\sigma})$ , if  $\dot{A} = (g_0)^{H(\dot{\sigma}||\dot{\beta})}$  and  $\dot{B} = (g_2)^{H(\dot{\sigma}||\dot{\beta})} \cdot (1 + \dot{\sigma}N) \bmod N^2$  both hold, then compute

$$\sigma = \frac{B/(A' \cdot A^{\dot{\beta}}) - 1 \bmod N'^2}{N'}, \quad m = C \oplus H_2(\sigma);$$

otherwise, output **reject** and terminate. If  $A = (g'_0)^{H'(\sigma||m)} \bmod N'^2$  and  $B = (g'_1)^{H'(\sigma||m)} \cdot (1 + \sigma N') \bmod N'^2$  both hold, output  $m$ ; otherwise, output **reject** and terminate.

Note that  $(H(\cdot), N, g_0, g_1, g_2)$  is the public key of the decryptor.

**Theorem 2.3.9** (CCA-O security). *In the random oracle model, scheme  $S_{\text{SUPRE}}$  is CCA-O-secure under the assumptions that the DDH problem over  $\mathbb{Z}_{N^2}^*$  is hard, and that the signature of knowledge is secure.*

*Proof.* . We show that if an algorithm  $\mathcal{A}$  exists and can break the CCA-O security of  $S_{\text{SUPRE}}$  with probability  $\epsilon$  in time  $t$ , there is another algorithm  $\mathcal{B}$  that uses  $\mathcal{A}$  to solve the DDH problem over  $\mathbb{Z}_{N^2}^*$ , i.e., on DDH input  $(N, g, g^u, g^v, T)$ ,  $\mathcal{B}$  decides if  $T = g^{uv}$  or not.

$\mathcal{B}$  interacts with  $\mathcal{A}$  in a CCA-O game as follows ( $\mathcal{B}$  simulates the challenger for  $\mathcal{A}$ ). In the following, we use starred letters  $(A^*, B^*, C^*, D^*, c^*, s^*)$  to refer to the challenge ciphertext corresponding to the target public key  $pk^*$ .

### Hash Oracles:

$\mathcal{O}_H$ : One oracle  $\mathcal{O}_H$  is corresponding to one hash function  $H(\cdot)$  which is a part of user's public key. As a result, there are many such oracles, and they are all constructed in the following method. On input  $(\sigma_i, m_i)$ ,  $\mathcal{B}$  first checks whether triple  $(\sigma_i, m_i, \alpha_i)$  is in Table  $T_H$ . If yes,  $\mathcal{B}$  responds  $\mathcal{A}$  with  $\alpha_i$ ; otherwise  $\mathcal{B}$  chooses a

random number  $\alpha_i \in \mathbb{Z}_{N^2}$ , responds  $\mathcal{A}$  with  $\alpha_i$ , and records  $(\sigma_i, m_i, \alpha_i)$  in Table  $T_H$ , where  $N$  is the corresponding safe-prime modulus and a part of user's public key.

$\mathcal{O}_{H_1}$ : On input  $\sigma_i$ ,  $\mathcal{B}$  first checks whether pair  $(\sigma_i, \beta_i)$  is in Table  $T_{H_1}$ . If yes,  $\mathcal{B}$  responds  $\mathcal{A}$  with  $\beta_i$ ; otherwise,  $\mathcal{B}$  chooses a random number  $\beta_i \in \{0, 1\}^{\lambda_1}$ , responds  $\mathcal{A}$  with  $\beta_i$ , and records  $(\sigma_i, \beta_i)$  in Table  $T_{H_1}$ .

$\mathcal{O}_{H_2}$ : On input  $\sigma_i$ ,  $\mathcal{B}$  first checks whether pair  $(\sigma_i, \gamma_i)$  is in Table  $T_{H_2}$ . If yes,  $\mathcal{B}$  responds  $\mathcal{A}$  with  $\gamma_i$ ; otherwise,  $\mathcal{B}$  chooses a random number  $\gamma_i \in \{0, 1\}^n$ , responds  $\mathcal{A}$  with  $\gamma_i$ , and records  $(\sigma_i, \gamma_i)$  in Table  $T_{H_2}$ .

$\mathcal{O}_{H_3}$ : On input  $(A_i, D_i, g_{i0}, g_{i2}, E_i, F_i, B_i, C_i)$ ,  $\mathcal{B}$  first checks whether tuple  $(A_i, D_i, g_{i0}, g_{i2}, E_i, F_i, B_i, C_i, \delta_i)$  is in Table  $T_{H_2}$ . If yes,  $\mathcal{B}$  responds  $\mathcal{A}$  with  $\delta_i$ ; otherwise,  $\mathcal{B}$  chooses a random number  $\delta_i \in \{0, 1\}^{\lambda_2}$ , responds  $\mathcal{A}$  with  $\delta_i$ , and records  $(A_i, D_i, g_{i0}, g_{i1}, E_i, F_i, B_i, C_i, \delta_i)$  in Table  $T_{H_3}$ .

### Phase 1:

$\mathcal{O}_{pk}$ : On input an index  $i$ ,  $\mathcal{B}$  decides whether  $pk_i$  is the target public key  $pk^*$ .

- If yes,  $\mathcal{B}$  sets  $N = \mathbf{N}$ ,  $H(\cdot) : \{0, 1\}^* \rightarrow \mathbb{Z}_{N^2}$ ,  $g_0 = \mathbf{g}$ ,  $g_1 = \mathbf{g}^u$ , and  $g_2 = \mathbf{g}^w$ , where  $w \in \mathbb{Z}_{N^2}$ . And then,  $\mathcal{B}$  records

$$(H(\cdot), N, g_0, g_1, g_2, \perp, \perp, \perp, \perp)$$

in Table  $T_k$ .

- Otherwise,  $\mathcal{B}$  runs KeyGen to get the public key  $(H(\cdot), N, g_0, g_1, g_2)$ , the private key  $(p', q')$ , and weak private key  $(a, b)$ ; and records

$$(H(\cdot), N, g_0, g_1, g_2, a, b, p', q')$$

in Table  $T_k$ .

Finally,  $\mathcal{B}$  returns  $(H(\cdot), N, g_0, g_1, g_2)$  to  $\mathcal{A}$  as  $pk_i$ .

$\mathcal{O}_{sk}$ : On input  $pk_X = (H_X(\cdot), N_X, g_{X0}, g_{X1}, g_{X2})$ ,  $\mathcal{B}$  checks whether  $pk_X$  is in  $T_k$ . If not,  $\mathcal{B}$  terminates. Otherwise, if  $pk_X$  is the guessed attacked public key,  $\mathcal{B}$  reports **failure** and aborts; otherwise,  $\mathcal{B}$  responds  $\mathcal{A}$  with corresponding  $(p'_X, q'_X)$ , and records  $pk_X$  in Table  $T_{sk}$ .

$\mathcal{O}_{rk}$ : On input  $(pk_X, pk_Y)$ ,  $\mathcal{B}$  checks whether  $pk_X$  and  $pk_Y$  are both in  $T_k$ . If not,  $\mathcal{B}$  terminates. Otherwise,  $\mathcal{B}$  checks whether  $(pk_X, pk_Y, rk_{X \rightarrow Y}^{(1)}, rk_{X \rightarrow Y}^{(2)})$  is in Table  $T_{rk}$ , or  $(pk_X, pk_Y, \beta_{X \rightarrow Y}, rk_{X \rightarrow Y}^{(1)}, rk_{X \rightarrow Y}^{(2)})$  is in Table  $T_{urk}$ , if it exists,  $\mathcal{B}$  returns  $(rk_{X \rightarrow Y}^{(1)}, rk_{X \rightarrow Y}^{(2)})$  to  $\mathcal{A}$ ; otherwise,

- if  $pk_X$  is in Table  $T_{sk}$  or  $pk_X$  is not the guessed attacked public key,  $\mathcal{B}$  responds  $\mathcal{A}$  with  $(rk_{X \rightarrow Y}^{(1)}, rk_{X \rightarrow Y}^{(2)}) \leftarrow \text{ReKeyGen}(sk_X, pk_Y)$ , and records  $(pk_X, pk_Y, rk_{X \rightarrow Y}^{(1)}, rk_{X \rightarrow Y}^{(2)})$  in Table  $T_{rk}$ ;
- if  $pk_X$  is the guessed attacked public key, and  $pk_Y$  is not in Table  $T_{sk}$ ,  $\mathcal{B}$  proceeds as follows.
  - Choose three random numbers  $\dot{\beta} \in \{0, 1\}^{\lambda_1}$ ,  $rk_{X \rightarrow Y}^{(2)} \in \mathbb{Z}_{N^2}$ ,  $\dot{\sigma} \in \mathbb{Z}_N$ .
  - Compute

$$\begin{aligned}
 r_{X \rightarrow Y} &= H_Y(\dot{\sigma} || \dot{\beta}), \\
 \dot{A} &= (g_{Y0})^{r_{X \rightarrow Y}} \bmod (N_Y)^2, \\
 \dot{B} &= (g_{Y1})^{r_{X \rightarrow Y}} \cdot (1 + \dot{\sigma} N_Y) \bmod (N_Y)^2, \\
 \dot{C} &= H_1(\dot{\sigma}) \oplus \dot{\beta},
 \end{aligned}$$

where  $(H_Y(\cdot), N_Y, g_{Y0}, g_{Y1}, g_{Y2}) = pk_Y$ .

- Set  $rk_{X \rightarrow Y}^{(1)} = (\dot{A}, \dot{B}, \dot{C})$ .
- Return  $(rk_{X \rightarrow Y}^{(1)}, rk_{X \rightarrow Y}^{(2)})$  to  $\mathcal{A}$ .

- Record  $(pk_X, pk_Y, \beta, rk_{X \rightarrow Y}^{(1)}, rk_{X \rightarrow Y}^{(2)})$  in Table  $T_{urk}$ .

If  $\mathcal{B}$  successfully responds the decryption queries with the re-encrypted ciphertexts that re-encrypted by the above re-encryption keys, these keys are indistinguishable from that in the real execution from the viewpoint of  $\mathcal{A}$ , due to the security property of scheme BCP03 with Fujisaki-Okamoto conversion.

- If  $pk_X$  is the guessed attacked public key, and  $pk_Y$  is in Table  $T_{sk}$ ,  $\mathcal{B}$  reports failure and aborts.

$\mathcal{O}_{re}$ : On input  $(pk_X, pk_Y, K)$ ,  $\mathcal{B}$  checks whether  $pk_X$  and  $pk_Y$  are both in Table  $T_k$ . If not,  $\mathcal{B}$  terminates. Otherwise,  $\mathcal{B}$  parses  $K = (A, B, C, D, c, s)$ , and checks whether  $c = H_3(A || D || g_{X0} || g_{X2} || (g_{X0})^s A^c || (g_{X2})^s D^c || (B || C))$ , where  $pk_X = (H_X(\cdot), N_X, g_{X0}, g_{X1}, g_{X2})$ , if not,  $\mathcal{B}$  outputs **reject** and terminates; otherwise,  $\mathcal{B}$  proceeds as follows.

- If  $pk_X$  is the guessed attacked public key, and  $pk_Y$  is in Table  $T_{sk}$ ,  $\mathcal{B}$  executes the following steps:

1. Set two empty sets,  $S_1$  and  $S_2$ .
2. Find all elements  $(\sigma_i, m_i, \alpha_i)$  in Table  $T_{H_X}$  such that  $A = (g_{X0})^{\alpha_i} \bmod (N_X)^2$ , and put them into Set  $S_1$ . If  $S_1 = \emptyset$ , output **reject** and terminate.

*This step makes this oracle distinguishable from the real execution when the adversary can guess the correct value of  $H_X(\sigma_i || m_i)$  without querying  $\mathcal{O}_{H_X}$ . The probability of this event is  $q_{H_X} / |\mathbb{Z}_{(N_X)^2}|$ , where  $q_{H_X}$  is the number of queries to  $\mathcal{O}_{H_X}$ .*

3. For every  $(\sigma_i, m_i, \alpha_i)$  in Set  $S_1$ , find all elements in Table  $T_{H_2}$  such that  $\sigma_j = \sigma_i$  and put them (i.e.,  $(\sigma_i, m_i, \alpha_i) || (\sigma_j, \gamma_j)$ 's) into Set  $S_2$ . If  $S_2 = \emptyset$ , output **reject** and terminate.

*This step makes this oracle distinguishable from the real execution when the adversary can guess the correct value of  $H_2(\sigma_i)$  without querying  $\mathcal{O}_{H_2}$ . The probability of this event is  $q_{H_2}/2^n$ , where  $q_{H_2}$  is the number of queries to  $\mathcal{O}_{H_2}$ .*

4. Check if  $(\sigma_i, m_i, \alpha_i) || (\sigma_j, \gamma_j)$  is in Set  $S_2$ , such that  $(g_{X1})^{\alpha_i} \cdot (1 + \sigma_i N_X) \bmod (N_X)^2 = B$  and  $\gamma_j \oplus m_i = C$ . If it does not exist or more than one exist, output **reject** and terminate.
5. Search  $(pk_X, pk_Y, \dot{\beta})$  in Table  $T_{f_{rk}}$ , if not, choose a random number from  $\{0, 1\}^{\lambda_1}$  as  $\dot{\beta}$ , and record  $(pk_X, pk_Y, \dot{\beta})$  in Table  $T_{f_{rk}}$ .
6. Choose a random number  $\dot{\sigma} \in \mathbb{Z}_N$ .
7. Compute

$$r_{X \rightarrow Y} = H_Y(\dot{\sigma} || \dot{\beta}),$$

$$\dot{A} = (g_{Y0})^{r_{X \rightarrow Y}},$$

$$\dot{B} = (g_{Y1})^{r_{X \rightarrow Y}} \cdot (1 + \dot{\sigma} N_Y) \bmod (N_Y)^2,$$

$$\dot{C} = H_1(\dot{\sigma}) \oplus \beta_{X \rightarrow Y},$$

where  $pk_Y = (H_Y(\cdot), N_Y, g_{Y0}, g_{Y1}, g_{Y2})$ .

8. Set  $rk_{X \rightarrow Y}^{(1)} = (\dot{A}, \dot{B}, \dot{C})$ .
  9. Return  $(A, (g_{X1})^{\alpha_i} \cdot (g_{X0})^{-\dot{\beta}}, B, C, rk_{X \rightarrow Y}^{(1)})$  to  $\mathcal{A}$ .
- Otherwise,  $\mathcal{B}$  calls oracle  $\mathcal{O}_{rk}$  to get the re-encryption key  $rk_{X \rightarrow Y}$ , and returns  $\text{ReEnc}(rk_{X \rightarrow Y}, K)$ .

$\mathcal{O}_{dec}$ : On input  $(pk_X, K)$ ,  $\mathcal{B}$  checks whether  $pk_X$  is in Table  $T_K$ , if not,  $\mathcal{B}$  terminates.

Otherwise,  $\mathcal{B}$  proceeds as follows:

- If  $pk_X$  is not the guessed attacked public key, then  $sk_X$  is known to  $\mathcal{B}$ , who responds  $\mathcal{A}$  with  $\text{Dec}(sk_X, K)$ .

- If  $pk_X$  is the guessed attacked public key and  $K = (A, B, C, D, c, s)$ ,  $\mathcal{B}$  checks whether  $c = H_3(A||D||g_{X0}||g_{X2}||(g_{X0})^s A^c||(g_{X2})^s D^c||(B||C))$ , if not,  $\mathcal{B}$  outputs **reject** and terminates; otherwise,  $\mathcal{B}$  does:

1. Set two empty sets  $S_1$  and  $S_2$ .
2. Find all elements  $(\sigma_i, m_i, \alpha_i)$  in Table  $T_{H_X}$  such that  $A = (g_{X0})^{\alpha_i} \bmod (N_X)^2$ , and put them into Set  $S_1$ . If  $S_1 = \emptyset$ , output **reject** and terminate.

*This step makes this oracle distinguishable from the real execution when the adversary can guess the correct value of  $H_X(\sigma_i||m_i)$  without querying  $\mathcal{O}_{H_X}$ . The probability of this event is  $q_{H_X}/|\mathbb{Z}_{(N_X)^2}|$ .*

3. For every  $(\sigma_i, m_i, \alpha_i)$  in Set  $S_1$ , find all elements in Table  $T_{H_2}$  such that  $\sigma_j = \sigma_i$  and put them (i.e.,  $(\sigma_i, m_i, \alpha_i)||(\sigma_j, \gamma_j)$ 's) into Set  $S_2$ . If  $S_2 = \emptyset$ , output **reject** and terminate.

*This step makes this oracle distinguishable from the real execution when the adversary can guess the correct value of  $H_2(\sigma_i)$  without querying  $\mathcal{O}_{H_2}$ . The probability of this event is  $q_{H_2}/2^n$ .*

4. Check if  $(\sigma_i, m_i, \alpha_i)||(\sigma_j, \gamma_j)$  is in Set  $S_2$ , such that  $(g_{X1})^{\alpha_i} \cdot (1 + \sigma_i N_X) \bmod (N_X)^2 = B$  and  $\gamma_j \oplus m_i = C$ . If none exists or more than one exist, output **reject** and terminate; otherwise, output  $m_i$ .
- If  $pk_X$  is the guessed attacked public key and  $K = (A, A', B, C, \dot{A}, \dot{B}, \dot{C})$ ,  $\mathcal{B}$  searches  $(pk_Y, pk_X, \star_1, \star_2, \star_3)$  in Table  $T_{urk}$ , such that  $\star_2 = (\dot{A}, \dot{B}, \dot{C})$ .

**If it does not exist**,  $\mathcal{B}$  proceeds as follows:

1. Set two empty sets  $S_1$  and  $S_2$ .
2. Find all elements  $(\sigma_i, m_i, \alpha_i)$  in Table  $T_{H_X}$  such that  $\dot{A} = (g_{X0})^{\alpha_i} \bmod (N_X)^2$ , and put them into Set  $S_1$ . If  $S_1 = \emptyset$ , output **reject** and terminate.

*This step makes this oracle distinguishable from the real execution when the adversary can guess the correct value of  $H_X(\sigma_i || m_i)$  without querying  $\mathcal{O}_{H_X}$ . The probability of this event is  $q_{H_X} / |\mathbb{Z}_{(N_X)^2}|$ .*

3. For every  $(\sigma_i, m_i, \alpha_i)$  in Set  $S_1$ , find all elements in Table  $T_{H_1}$  such that  $\sigma_j = \sigma_i$  and put them (i.e.,  $(\sigma_i, m_i, \alpha_i) || (\sigma_j, \beta_j)$ 's) into Set  $S_2$ . If  $S_2 = \emptyset$ , output **reject** and terminate.

*This step makes this oracle distinguishable from the real execution when the adversary can guess the correct value of  $H_1(\sigma_i)$  without querying  $\mathcal{O}_{H_1}$ . The probability of this event is  $q_{H_1} / 2^{\lambda_1}$ , where  $q_{H_1}$  is the number of queries to  $\mathcal{O}_{H_1}$ .*

4. Check if  $(\sigma_i, m_i, \alpha_i) || (\sigma_j, \beta_j)$  is in Set  $S_2$ , such that  $\dot{B} = (g_{X1})^{\alpha_i} \cdot (1 + \sigma_i N_X) \bmod (N_X)^2$  and  $\beta_j \oplus m_i = \dot{C}$ . If none exists or more than one exist, output **reject** and terminate.
5. Compute  $\sigma = \frac{(B / (A' \cdot A^{m_i}) - 1 \bmod (N_Y)^2)}{N_Y}$ ,  $m = C \oplus H_2(\sigma)$ . If  $A = (g_{Y0})^{H_Y(\sigma || m)} \bmod (N_Y)^2$ , output  $m$ , where  $pk_Y = (H_Y(\cdot), N_Y, g_{Y0}, g_{Y1}, g_{Y2})$  is the corresponding delegator's public key; otherwise, output **reject** and terminate.

**If it does exist**,  $\mathcal{B}$  checks  $A' \stackrel{?}{=} A^{\star 3}$ . If not, output **reject** and terminate; otherwise,  $\mathcal{B}$  proceeds as follows:

1. Set two empty sets  $S_1$  and  $S_2$ .
2. Find all elements  $(\sigma_i, m_i, \alpha_i)$  in Table  $T_{H_Y}$  such that  $A = (g_{Y0})^{\alpha_i} \bmod (N_Y)^2$ , and put them into Set  $S_1$ . If  $S_1 = \emptyset$ , output **reject** and terminate.

*This step makes this oracle distinguishable from the real execution when the adversary can guess the correct value of  $H_Y(\sigma_i || m_i)$  without querying  $\mathcal{O}_{H_Y}$ . The probability of this event is  $q_{H_Y} / |\mathbb{Z}_{(N_Y)^2}|$ , where*

$q_{H_Y}$  is the number of queries to  $\mathcal{O}_{H_Y}$ .

3. For every  $(\sigma_i, m_i, \alpha_i)$  in Set  $S_1$ , find all elements in Table  $T_{H_2}$  like  $\sigma_j = \sigma_i$  and put them (i.e.,  $(\sigma_i, m_i, \alpha_i) || (\sigma_j, \gamma_j)$ 's) into Set  $S_2$ . If  $S_2 = \emptyset$ , output **reject** and terminate.

*This step makes this oracle distinguishable from the real execution when the adversary can guess the correct value of  $H_2(\sigma_i)$  without querying  $\mathcal{O}_{H_2}$ . The probability of this event is  $q_{H_2}/2^n$ .*

4. Check if  $(\sigma_i, m_i, \alpha_i) || (\sigma_j, \gamma_j)$  is in Set  $S_2$ , such that  $(g_{Y1})^{\alpha_i} \cdot (1 + \sigma_i N_Y) \bmod (N_Y)^2 = B$  and  $\gamma_j \oplus m_i = C$ . If none exists or more than one exist, output **reject** and terminate; otherwise, output  $m_i$ .

**Challenge:** At some point,  $\mathcal{A}$  outputs a challenge tuple  $(pk^*, m_0, m_1)$ . If  $pk^*$  is not the public key  $\mathcal{B}$  guessed in oracle  $\mathcal{O}_{pk}$ ,  $\mathcal{B}$  reports **failure** and aborts. Otherwise,  $\mathcal{B}$  chooses random  $d \in \{0, 1\}$ ,  $\sigma \in \mathbb{Z}_N$  and sets:

$$A^* = \mathbf{g}^{\mathbf{v}} \bmod \mathbf{N}^2, \quad B^* = \mathbf{T}(1 + m_d \mathbf{N}) \bmod \mathbf{N}^2,$$

$$C^* = H_2(\sigma) \oplus m_d, \quad D^* = (\mathbf{g}^{\mathbf{v}})^w \bmod \mathbf{N}^2.$$

And then  $\mathcal{B}$  chooses two random numbers  $c^* \in \{0, 1\}^{\lambda_2}$ ,  $s^* \in \{0, \dots, 2^{L(\mathbf{N}^2) + \lambda_2} - 1\}$ , computes  $E^* = (\mathbf{g})^{s^*} A^{*c^*} \bmod \mathbf{N}^2$  and  $F^* = (\mathbf{g}^{\mathbf{u}})^{s^*} D^{*c^*} \bmod \mathbf{N}^2$ , and checks whether  $(A^*, D^*, \mathbf{g}^{\mathbf{u}}, \mathbf{g}^w, E^*, F^*, B^*, C^*, \star)$  is in Table  $T_{H_3}$ . If yes,  $\mathcal{B}$  reports **failure** and aborts; otherwise,  $\mathcal{B}$  outputs  $(A^*, B^*, C^*, D^*, c^*, s^*)$ , and records

$$(A^*, D^*, \mathbf{g}^{\mathbf{u}}, \mathbf{g}^w, E^*, F^*, B^*, C^*, c^*)$$

in Table  $T_{H_3}$ .

### Phase 2:

$\mathcal{O}_{pk}$ :  $\mathcal{B}$  responds as in Phase 1.



$\mathcal{O}_{sk}$ : On input  $pk_i$ , if  $pk_i = pk^*$ , or  $(pk^*, pk_i)$  is in Table  $T_{rk}$ , then  $\mathcal{B}$  terminates.

Otherwise,  $\mathcal{B}$  responds as in Phase 1.

$\mathcal{O}_{rk}$ : On input  $(pk_i, pk_j)$ , if  $pk_i = pk^*$ , and  $pk_j$  is in Table  $T_{sk}$ ,  $\mathcal{B}$  terminates. Otherwise,  $\mathcal{B}$  responds as in Phase 1.

$\mathcal{O}_{re}$ : On input  $(pk_i, pk_j, K)$ , if  $(pk_i, K) = (pk^*, K^*)$  and  $pk_j$  is in Table  $T_{sk}$ ,  $\mathcal{B}$  terminates. Otherwise,  $\mathcal{B}$  responds as in Phase 1, except when  $pk_i = pk^*$  and  $(A, B, C, D, c, s) = (A^*, B^*, C^*, D^*, c^*, s^*)$ ,  $\mathcal{B}$  should record the result  $(pk_j, A', C, D, \dot{A}, \dot{C}, \dot{D})$  in Table  $T_{der}$ , where the derivatives of the challenge ciphertext are recorded.

$\mathcal{O}_{dec}$ : On input  $(pk_i, K)$ , if  $(pk_i, K) = (pk^*, K^*)$ , or  $(pk_i, K)$  is in  $T_{der}$ , or  $K = \text{ReEnc}(\mathcal{O}_{rk}(pk^*, pk_i), K^*)$ , then  $\mathcal{B}$  terminates. Otherwise,  $\mathcal{B}$  responds as in Phase 1.

**Guess:** Finally, the adversary  $\mathcal{A}$  outputs a guess  $d' \in \{0, 1\}$ . If  $d = d'$ , then  $\mathcal{B}$  outputs 1 (i.e., DDH instance), otherwise,  $\mathcal{B}$  outputs 0 (i.e., not a DDH instance).

Firstly, we analyze the probability of that  $\mathcal{B}$  does not output `failure`, which contains the following two events:

1.  $\mathcal{B}$  guesses the right attacked public key.
2. The record  $(A^*, D^*, \mathbf{g}^u, \mathbf{g}^w, E^*, F^*, B^*, C^*, \star)$  is not in table  $T_{H_3}$  before challenge phase.

Suppose  $\mathcal{A}$  makes a total of  $q_{pk}$  queries to public key generation oracle,  $q_{rk}$  queries to re-encryption key generation oracle,  $q_{de}$  queries to decryption oracle,  $q_H$  queries to  $H$  hash function oracle,  $q_{H_1}$  queries to  $H_1$  hash function oracle,  $q_{H_2}$  queries to  $H_2$  hash function oracle, and  $q_{H_3}$  queries to  $H_3$  hash function oracle.

The probabilities of the above two events are  $1/q_{pk}$  and  $1 - (q_{H_3} + 1)/2^{\lambda_2}$ , respectively. Therefore, the probability that  $\mathcal{B}$  does not output `failure` during the simulation

is  $(1 - (q_{H_3} + 1)/2^{\lambda_2})/q_{pk}$ .

Secondly, oracles  $\mathcal{O}_{re}$  and  $\mathcal{O}_{dec}$  are indistinguishable from the corresponding real executions with probabilities at least

$$\begin{aligned} & \left( \frac{1 + q_{max} + (q_{max})^2}{(1 + q_{max})^2} + \frac{q_{max}}{(1 + q_{max})^2} \left( 1 - \frac{q_{H_X}}{|\mathbb{Z}_{N_{mX}}|} \right) \left( 1 - \frac{q_{H_1}}{2^{\lambda_1}} \right) \right)^{q_{re}} \\ & > \left( \frac{1 + q_{max} + (q_{max})^2}{(1 + q_{max})^2} \right)^{q_{re}} \end{aligned}$$

and

$$\left( \frac{q_{max}}{1 + q_{max}} + \frac{1}{1 + q_{max}} \left( 1 - \frac{q_{H_X}}{|\mathbb{Z}_{N_{mX}}|} \right) (1 - q_2) \right)^{q_{de}} > \left( \frac{q_{max}}{1 + q_{max}} \right)^{q_{de}},$$

respectively, where  $q_{H_X}$  is the amount of queries to the same kind of oracle  $\mathcal{O}_H$ ,  $N_{mX}$  is the largest number among users' public key  $N$ 's,  $|\mathbb{Z}_{N_{mX}}|$  is the size of  $\mathbb{Z}_{N_{mX}}$ , and  $q_2 = \max(\frac{q_{H_1}}{2^{\lambda_1}}, \frac{q_{H_2}}{2^n})$ .

Finally, in the re-encryption oracle, we assume that the signature of knowledge is secure, hence, we should minus the probability of breaking the signature of knowledge,  $\xi$ .

As a result,  $\mathcal{B}$  has an advantage that is at least

$$\epsilon \cdot \frac{(1 - (q_{H_3} + 1)/2^{\lambda_2}) \cdot (1 + q_{max} + (q_{max})^2)^{q_{re}} \cdot (q_{max})^{q_{de}}}{q_{pk} \cdot (1 + q_{max})^{2q_{re} + q_{de}}} - \xi$$

and its running time is at most

$$t + O(3q_{pk} + (7 + q_H)q_{rk} + (5 + q_H)q_{de})t_e,$$

where  $t_e$  is the time of computing one exponentiation in a cyclic group of quadratic residues modulo. We here only consider the exponentiation computation.

This completes the proof.  $\square$

**Theorem 2.3.10** (CCA-R security). *In the random oracle model, scheme  $S_{\text{SUPRE}}$  is CCA-*

*R-secure under the assumptions that the DDH problem over  $\mathbb{Z}_{N^2}^*$  is hard, and that the signature of knowledge is secure.*

We can use a similar method in the proof of Theorem 2.3.9 to prove the above theorem. The main difference is that the public key of the target user is  $(N = \mathbf{N}, g_0 = g, g_1 = \mathbf{g}^w, g_2 = \mathbf{g}^u)$  not  $(N = \mathbf{N}, g_0 = g, g_1 = \mathbf{g}^u, g_2 = \mathbf{g}^w)$ . This change could help us to generate any re-encryption key from the target user to any other user.

We leave this proof to the reader for practice.

## 2.4 Notes

In this chapter, we give a comprehensive introduction of security definitions of proxy re-cryptography. We also introduce one of our PRS schemes and one of PRE schemes, which are published at Indocrypt 2007 and PKC 2009, respectively. We refer to [11, 193, 252] and [9, 10, 130, 194, 195, 249, 255, 283] for more PRS schemes and PRE schemes, respectively.

There are some interesting problems in proxy re-cryptography still remaining open. For instance,

- How to design proxy re-signature (re-encryption) with multi-usability, unidirectionality, and constant size.
- How to design pairing-free proxy re-signature (re-encryption) with collusion resistance, multi-usability, bidirectionality, and constant size.
- How to design generic constructions of PRE with some properties from the public key encryption with the specific properties, like the generic construction from ID-based (attribute-based or certificateless) encryption to ID-based (attribute-based or certificateless) PRE.

## Chapter 3

# Attribute-Based Cryptography

### 3.1 Introduction

With the development of communication networks, there is a trend for users to store their sensitive data on the Internet. To distribute a message to a specific set of users, a trivial method is to encrypt it under each user's public key or identity in traditional cryptosystem [34, 37, 87, 89, 90]. As expected, ciphertext size and computational cost of encryption/decryption algorithms are linear with the number of receivers. Therefore, it is less attractive or even intolerable when the number of receivers is large. Indeed, in most cases, the qualified receivers share some common attributes, such as working location, gender, and age range.

For this reason, Sahai and Waters took the first step to solve the problem and introduced the concept of attribute-based encryption. In an attribute-based encryption mechanism (ABE), user's keys and ciphertexts are labeled with sets of descriptive attributes and only when the attributes of the ciphertext match those of the user's key, the corresponding ciphertext can be decrypted. The proposed scheme allows for decryption when a threshold number  $k$  is less than the size of the overlap between a ciphertext

and a private key. The attribute-based cryptosystems that each ciphertext is labeled by the encryptor with a set of descriptive attributes and each private key is associated with an access structure that specifies the types of ciphertexts the key can decrypt, are named key-policy attribute-based encryption (KPABE). While in a ciphertext-policy attribute-based encryption scheme (CPABE), each user is identified by an attribute set and receives the private keys corresponding to those attributes from the authority. The sender who aims to distribute a message will construct an access policy associated to the ciphertexts by connecting the attributes with OR, AND, and threshold gates.

The remainder of this chapter is organized as follows. Some universal definitions are introduced in Section 3.2. Then, a bounded ciphertext-policy encryption scheme, a multi-authority encryption scheme, an interval encryption scheme, and a fuzzy identity-based signature scheme are proposed in Section 3.3, 3.4, 3.5, and 3.6, respectively. Finally, a brief conclusion is given and some future works are suggested in Section 3.7.

## 3.2 Universal Definitions

**Definition 3.2.1** (Access Structure [20]). Let  $\{P_1, P_2, \dots, P_n\}$  be a set of parties. A collection  $\mathbb{A} \subseteq 2^{\{P_1, P_2, \dots, P_n\}}$  is monotone, for  $\forall B, C$ , if  $B \in \mathbb{A}$  and  $B \subseteq C$ , then  $C \in \mathbb{A}$ . An access structure (monotone access structure) is a collection (monotone collection)  $\mathbb{A}$  of nonempty subsets of  $\{P_1, P_2, \dots, P_n\}$ , i.e.,  $\mathbb{A} \subseteq 2^{\{P_1, P_2, \dots, P_n\}} \setminus \{\emptyset\}$ . The sets in  $\mathbb{A}$  are called the authorized sets, and the sets not in  $\mathbb{A}$  are called the unauthorized sets.

In ABE, the access structure  $\mathbb{A}$  contains the authorized sets of attributes. It is shown [20] that any monotone access structure can be realized by a linear secret sharing scheme.

**Definition 3.2.2** (Linear Secret-Sharing Schemes (LSSS) [291]). A secret sharing scheme  $\Pi$  over a set of parties  $\mathbb{P}$  is called linear (over  $\mathbb{Z}_p$ ) if

1. The shares for each party form a vector over  $\mathbb{Z}_p$ .
2. There exists a matrix  $A$  called the share-generating matrix for  $\Pi$ . The matrix  $A$  has  $l$  rows and  $n$  columns. For  $i = 1, \dots, l$ , the  $i^{th}$  row of  $A$  is labeled by a party  $\rho(i)$  ( $\rho$  is a function from  $\{1, \dots, l\}$  to  $\mathbb{P}$ ). When we consider the column vector  $\vec{v} = (s, r_2, \dots, r_n)$ , where  $s \in \mathbb{Z}_p$  is the secret to be shared and  $r_2, \dots, r_n \in \mathbb{Z}_p$  are randomly chosen,  $A\vec{v}$  is the vector of  $l$  shares of the secret  $s$  according to  $\Pi$ . The share  $(A\vec{v})_i$  belongs to party  $\rho(i)$ .

It is shown that each linear secret-sharing scheme according to the above definition also enjoys the linear reconstruction property [20], defined as follows: Suppose that  $\Pi$  is an LSSS for access structure  $\mathbb{A}$ . Let  $S \in \mathbb{A}$  be an authorized set, and let  $I \subset \{1, 2, \dots, l\}$  be defined as  $I = \{i : \rho(i) \in S\}$ . There exist constants  $\{\omega_i \in \mathbb{Z}_p\}_{i \in I}$  such that if  $\{\lambda_i\}$  are valid shares of any secret  $s$  according to  $\Pi$ , then  $\sum_{i \in I} \omega_i \lambda_i = s$ . Furthermore, these constants  $\{\omega_i\}$  can be found in time polynomial to the size of the share-generating matrix  $A$ . For any unauthorized set, no such constants exist. In expressive CPABE systems, LSSS matrix  $(A, \rho)$  is always used to express an access policy associated to a ciphertext.

### 3.3 Bounded Ciphertext-Policy Encryption Schemes

The concept of attribute-based encryption (ABE) was introduced by Sahai and Waters [241]. In their scheme, the secret key is associated with an attribute set, and the ciphertext is also associated with another set of attributes. The decryption is only successful while these two sets overlap more than a preset threshold. Later, Goyal *et al.* [129] further classified ABE into two categories: ciphertext-policy (CP) ABE and key-policy (KP) ABE. While KPABE (the user's secret key represents an access policy of attributes) is well developed by the subsequent research [129, 223], how to design an efficient and secure CPABE (the ciphertext represents an access policy chosen by the

sender) remains open [27, 84, 127].

The first CPABE scheme was proposed by Bethencourt et al. [27] It allows users to encrypt a message under an expression consisting of threshold gates between attributes (called a fine-grained access structure). However, it only has a security argument in the generic group model and the random oracle model. Later, Cheung and Newport [84] gave a provably secure CPABE in the standard model. Their scheme supports an access policy with "AND" gate on positive and negative attributes but can not resist collusion attack while extending to the threshold gate. Recently, a bounded CPABE scheme, supporting fine-grained access policy, was proposed in [127].

Waters [291] presented several CPABE schemes. The construction of the first scheme is very elegant, and the security can be reduced to decisional  $q$ -Bilinear Diffie-Hellman-Exponent (BDHE) problem. The ciphertext size only linearly increases with the number of the attributes presented in the access structure. Another scheme, based on DBDH assumption, is less efficient, since the ciphertext size is restricted by the length and the width of the matrix, which is dependent on the size of the access structure.

### Overview of Schemes

Ciphertext-policy attribute-based encryption (CPABE) cryptosystems are widely used to realize the sensitive user data sharing and access control on the Internet. For example, "(AGE > 25) AND (CS)" represents the restrictions for a qualified user who at least holds the secret keys representing age more than 25 and the ownership of a computer science degree; three attributes (DB, OS, DM) connecting by a "two out of three threshold gate" restrict the decryption only to be successful when a user at least registers two courses from the list (Database, Operating System, and Discrete Mathematics). As in Figure 3.1, a sender will store the ciphertext encrypted under the access policy in the server  $S_1$ . The users  $U_2$  and  $U_3$  have the access right to the message since they both have two attributes from the attribute set, while  $U_1$  does not, since he/she only has

one attribute. This is an efficient and convenient approach for a user to broadcast his message to the others in practice. Therefore, we mainly focus on developing a more efficient CPABE scheme.

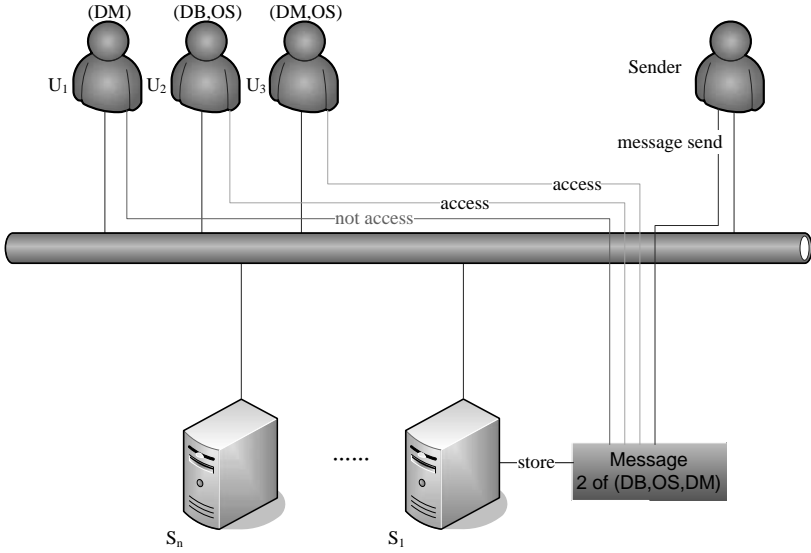


Figure 3.1: A sample for CPABE system

Considering the security proof of an ABE scheme, more flexible access policy adopted on the sender's side makes the simulation more difficult. Thus, restricting the size of the access policy is necessary for designing a CPABE which reduces the security to an assumption in number theory. Goyal *et al.* introduced a new bounded CPABE (BCPABE), in which the encryption access tree must be limited by two properties: the maximum height and the maximum cardinality of non-leaf nodes. For example, the access tree (to the left of Figure 3.2) with height 2 has the maximum cardinality 3.

In [127], a bounded CPABE scheme, secure in the standard model, was proposed to support a bounded size access tree with threshold gates as its non-leaf nodes. At the beginning, the system manager pre-sets two bounds  $(d, c)$  and a unique  $(d, c)$ -universal



access tree  $\mathcal{T}_u$  (Figure 3.3). Later, the system manager will publish the public parameter and generate users' secret keys according to this universal access tree. If a sender wants to distribute his message, it requires him to first convert a  $(d, c)$ -bounded access tree  $\mathcal{T}$  (the left in Figure 3.2) into its  $(d, c)$ -normal form<sup>1</sup>  $\mathcal{T}_n$  (the right in Figure 3.2), and then to complete the encryption procedure according to a map constructed from a  $(d, c)$ -normal form access tree  $\mathcal{T}_n$  to the  $(d, c)$ -universal access tree  $\mathcal{T}_u$ .

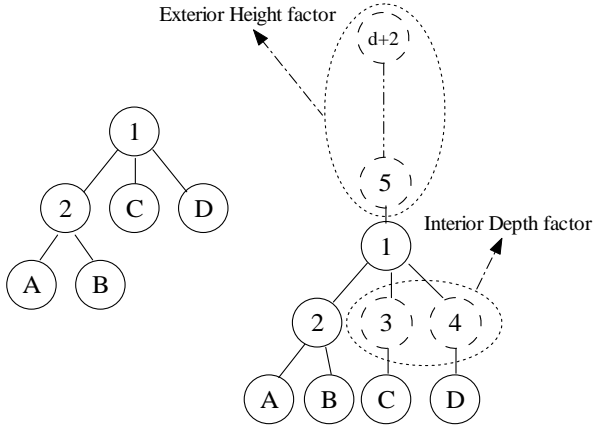


Figure 3.2: A conversion from  $(d, c)$ -bounded access tree  $\mathcal{T}$  to its  $(d, c)$ -normal form  $\mathcal{T}_n$

In the  $(d, c)$ -universal access tree  $\mathcal{T}_u$ , only leaf nodes of depth  $d$  are associated with real attributes. This setting leads to the fact that users have to construct a map from  $\mathcal{T}_n$  to  $\mathcal{T}_u$  in order to ensure that the leaf nodes in  $\mathcal{T}_n$  share same real attributes with their corresponding leaf nodes in  $\mathcal{T}_u$ . However, this conversion of normal form actually expands the original tree  $\mathcal{T}$  with a great many non-leaf nodes which leads to a boost on computational cost. We conclude that the expanded size is mainly due to two factors called exterior height factor and interior depth factor (refer to Figure 3.2). Exterior height factor correlates with the height  $h$  of a  $(d, c)$ -bounded access tree  $\mathcal{T}$ .  $d - h$  nodes must be added in order to expand the height of its normal form to  $d$ . Goyal *et*

<sup>1</sup>The definition of normal form can be looked up in Section 2.4

al. provide a method to eliminate this factor by constructing multiple parallel schemes with different-sized universal access trees, though it is very inefficient. Interior depth factor is the relative depth between leaf nodes in  $\mathcal{T}$ . Non-leaf nodes must be added in order to make the leaf nodes all at the same depth (which is the deepest level). In order to eliminate both factors, we neglect the interim step, i.e., converting the access tree to normal form, and directly map the  $(d, c)$ -bounded access tree selected by the sender to the  $(d, c)$ -universal tree. In other words, the redundant steps which pull all the leaf nodes to the deepest level by adding non-leaf nodes are skipped and thus the computational cost is reduced.

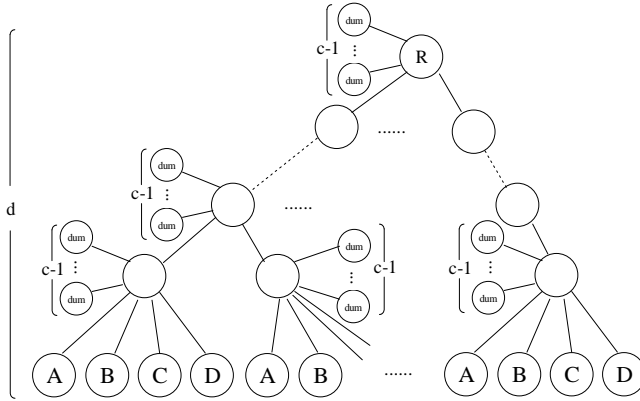


Figure 3.3: A  $(d, c)$ -universal access tree  $\mathcal{T}_u$  in GJPS

As mentioned in GJPS, to construct a more efficient BCPABE scheme based on some assumptions in number theory is an important open problem. We will provide an affirmative answer to this problem.

This section will present a bounded CPABE scheme  $\text{BCP}_1$  which is more efficient than the previous works in GJPS. The security of  $\text{BCP}_1$  can be reduced to the Decisional Bilinear Diffie–Hellman assumption in the standard model. Different from GJPS, the computational cost of encryption and decryption in our scheme are largely reduced

since we skip redundant steps, meanwhile the ciphertext size is shorter. Nevertheless, as a tradeoff, we demonstrate that the spacial cost, like the size of public parameter and secret keys, increases but be less than twice of the counterpart of GJPS. Furthermore, we propose a provably secure BCPABE scheme  $\text{BCP}_2$  in the standard model under a chosen ciphertext secure notion by adopting a one-time signature technique. There are two approaches which could be used for extending  $\text{BCP}_1$  to  $\text{BCP}_2$ . These two methods make a tradeoff between ciphertext size and the size of public/secret parameters.

For illustration, we take a concrete example to explain the primitive idea used to reduce the computational cost of encryption/decryption algorithms.

Consider a BCPABE scheme setup with bounds  $(d, c)^2$  and an encryption under a  $(d, c)$ -bounded access tree  $\mathcal{T}$  shown in the left of Figure 3.2. The threshold values of nodes 1 and 2 are both set to be 2, and nodes  $A, B, C, D$  represent four different real attributes. Now, we will show how the encryption algorithm of GJPS differs from ours.

In GJPS, to encrypt under  $\mathcal{T}$ , a user will first convert  $\mathcal{T}$  to its normal form access tree  $\mathcal{T}_n$  (shown in the right of Figure 3.2). The threshold values of nodes 3, 4, 5,  $\dots$ ,  $d+2$  in  $\mathcal{T}_n$  are all set to be 1. Assume the computational cost of a single node is  $T$ , the total cost is  $(4+c-2+c-2+d \times (c-1)) \cdot T$  where exterior factor takes  $(d-2) \cdot (c-1) \cdot T$  and interior factor takes  $2 \cdot (c-1) \cdot T$ . Likewise, if a user has a secret key associated with attributes  $(C, D)$ , he will at least expend  $(1+2+d-2) \times c = (d+1) \times c$  paring computation. It is obvious that the larger the two initial parameters  $(d, c)$  are set to be, the more cost a user will spend on encryption and decryption.

In contrast, our scheme defines a map from  $\mathcal{T}$  to  $\mathcal{T}_u$  and the total cost on encryption takes  $(4+c-2+c-2) \cdot T$ . The user with attributes  $(C, D)$  will expend only  $c$  paring computation. Therefore, our scheme saves  $d \times (c-1) \cdot T$  on encryption and at least  $dc$  times paring computation on decryption. The cost is even irrelevant with the initial parameter  $d$ . The comparison of general case can be found in Table 3.1 in the following

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<sup>2</sup> $d \geq 2, c \geq 3$

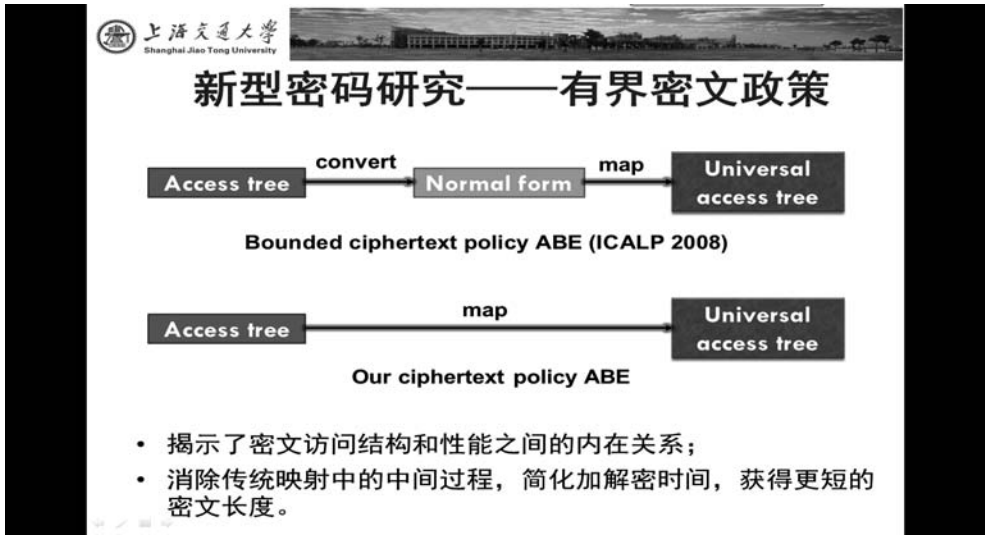


Figure 3.4: The comparison between GJPS's scheme in ICALP 2008 and our scheme

subsections. Figure 3.4 gives an overall comparison between GJPS's scheme in ICALP 2008 and our scheme.

### 3.3.1 Definitions

**Decisional Bilinear Diffie–Hellman Problem.** An algorithm  $\mathcal{S}$  is an  $\varepsilon'$ -solver of the DBDH problem if it distinguishes with probability at least  $1/2 + \varepsilon'$  between the two following probability distributions:

$$\mathcal{D}_{bdh} = (g, g^a, g^b, g^c, e(g, g)^{abc}), \text{ where } a, b, c \text{ are chosen randomly in } Z_p,$$

$\mathcal{D}_{rand} = (g, g^a, g^b, g^c, Z)$ , where  $a, b, c$  are chosen randomly in  $Z_p$  and  $Z$  is chosen randomly in  $G_T$ .

**Definition 3.3.1.** The DBDH assumption holds in  $G$  and  $G_T$  if there is no probabilistic polynomial-time  $\varepsilon'$ -solver of the DBDH problem for non-negligible value  $\varepsilon'$ .

**One-Time Signature.** A one-time signature scheme  $\text{ots}$  [36, 63] consists of three algorithms ( $\text{ots.KGen}$ ,  $\text{ots.Sig}$ ,  $\text{ots.Ver}$ ).  $\text{ots.KGen}(1^k) \rightarrow (sk, vk)$  is the key generation algorithm, which outputs a secret key  $sk$  and a public verification key  $vk$ .  $\text{ots.Sig}(sk, m) \rightarrow \sigma$  is the sign algorithm which takes the secret key  $sk$  and a message  $m$  as its input, and outputs a signature  $\sigma$ . Finally, the verification algorithm  $\text{ots.Ver}(\sigma, m, vk) \rightarrow 0 \text{ or } 1$  takes the signature  $\sigma$ , a message  $m$ , and a public veri-

fication key  $vk$  as its input, and outputs 1 if the signature is valid; 0 otherwise.

Concerning the security issue, an adversary first receives a public verification key  $vk$  generated from  $\text{ots.KGen}(1^k)$ . He then makes at most one signature query for message  $m$  of his choice, and obtains as an answer the valid signature  $\text{ots.Sig}(sk, m) \rightarrow \sigma$ . Finally, he outputs a pair  $(m', \sigma')$ . The adversary succeeds if  $(m', \sigma') \neq (m, \sigma)$  and  $\text{ots.Ver}(\sigma', m', vk) \rightarrow 1$ .

A one-time signature scheme  $\text{ots}$  is  $\varepsilon_{\text{ots}}$ -secure if every polynomial-time adversary against  $\text{ots}$  has a success probability bounded by  $\varepsilon_{\text{ots}}$ .

Several definitions and notions are given below.

**Attribute Set:**  $n$  real attributes, indexed from 1 to  $n$ . Any attribute set  $\gamma \subseteq \{1, \dots, n\}$ .  $c - 1$  dummy attributes, indexed from  $n + 1$  to  $n + c - 1$ .

**Access Tree:** Let  $\mathcal{T}$  represent an access tree with its root  $r$ . Each non-leaf node  $x$  can be seen as a threshold gate with threshold value  $k_x$ . If  $x$  has  $c_x$  child nodes, it is required that  $0 < k_x \leq c_x$ . If  $x$  is a leaf node of the access tree, it is associated with a single attribute, denoted as  $\text{att}(x)$ .

We fix the root of an access tree to be at depth 0. Let  $\Sigma_{\mathcal{T}}$  denote the set of all the non-leaf nodes, and  $\Theta_{\mathcal{T}}$  denote all the leaf nodes. Let  $\text{p}(x)$  denote the parent of node  $x$ . For each non-leaf node  $x$ , we define an order among  $x$ 's child nodes, that is, every child node  $z$  is numbered from 1 to  $c_x$ .  $\text{index}(z)$  returns such a number associated to node  $z$ . For simplicity, if  $z$  is a leaf node, we let  $\text{att}(z) = \text{index}(z)$ .

**Satisfying an Access Tree:** Let  $\mathcal{T}$  be an access tree with root  $r$ .  $\mathcal{T}_x$  is a subtree rooted at a node  $x$  in  $\mathcal{T}$ . If an attribute set  $\gamma$  satisfies the access tree  $\mathcal{T}_x$ , we denote  $\mathcal{T}_x(\gamma) = 1$ . If  $x$  is a non-leaf node, evaluate  $\mathcal{T}_z(\gamma)$  for all children  $z$  of node  $x$ .  $\mathcal{T}_x(\gamma)$  returns 1 if and only if at least  $k_x$  children of  $x$  return 1. If  $x$  is a leaf node, then  $\mathcal{T}_x(\gamma) = 1$  iff  $\text{att}(x) \in \gamma$ . If an access tree rooted at  $x$  is a  $\gamma$ -satisfied tree, then we call the node  $x$  as a  $\gamma$ -satisfied node. Suppose  $\mathcal{T}$  is a  $\gamma$ -satisfied tree, we call a subtree of  $\mathcal{T}$  with root  $r$  as a  $\gamma$ -satisfied non-redundant tree, if the cardinality of each non-leaf node

$x$  is equal to  $x$ 's threshold value in  $\mathcal{T}$  and every node is a qualified node. Let  $\hat{\mathcal{T}}$  denote the  $\gamma$ -satisfied nonredundant tree with minimum non-leaf nodes.<sup>3</sup>

**Universal Access Tree:** Here, we describe a universal access tree (Figure 3.5) with two input parameters  $d$  and  $c$ . First, we define a complete  $c$ -ary tree  $\mathcal{T}'$  of height  $d - 1$ , where each node has a threshold value  $c$ . Next,  $c - 1$  new leaf nodes named “dummy nodes” representing  $c - 1$  dummy attributes and  $n$  new leaf nodes named “real nodes” representing  $n$  real attributes are attached to each node in  $\mathcal{T}'$ . The resultant tree  $\mathcal{T}_u$  is called a  $(d, c)$ -universal access tree. Let each node  $x$  except root has an index related with its parent node and  $\text{att}(x) = \text{index}(x)$  if  $x$  is a leaf node of a universal access tree. Here, for all the child nodes of one non-leaf node  $x$  in  $\mathcal{T}$ , real nodes and dummy nodes will take indexes from  $\{1, \dots, n\}$  and  $\{n + 1, \dots, n + c - 1\}$ , respectively, and other non-leaf nodes will take indexes from  $\{n + c, \dots, n + 2c - 1\}$ .

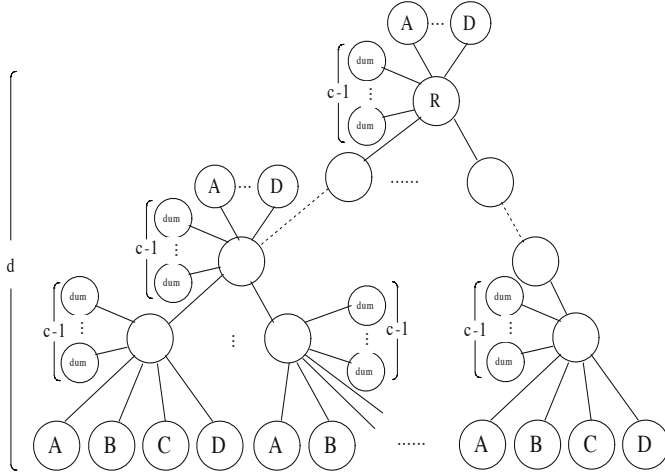


Figure 3.5: Modified universal tree

**Bounded Access Tree:** We call that  $\mathcal{T}$  a  $(d, c)$ -bounded access tree if its height  $d' \leq d$  and each non-leaf node  $x$  in  $\mathcal{T}$  has at most  $c$  non-leaf child nodes.

<sup>3</sup>The decryption cost depends on the non-leaf nodes in the  $\gamma$ -satisfied nonredundant tree the decryptor chooses.

**Normal Form:** Consider a  $(d, c)$ -bounded access tree  $\mathcal{T}_n$ .  $\mathcal{T}_n$  exhibits the  $(d, c)$ -normal form if (a) its height  $d' = d$ , and (b) all the leaf nodes are at depth  $d$ . Any  $(d, c)$ -bounded access tree  $\mathcal{T}$  can be converted to its normal form without modifying its satisfying logic. (This is a special technique used in GJPS.)

**Map between Access Trees:** The map, constructed from a  $(d, c)$ -bounded access tree  $\mathcal{T}$  to  $(d, c)$ -universal access tree  $\mathcal{T}_u$ , is defined in the following way in a top-down manner. First, the root of  $\mathcal{T}$  is mapped to the root of  $\mathcal{T}_u$ . Now, suppose that  $x'$  in  $\mathcal{T}$  is mapped to  $x$  in  $\mathcal{T}_u$ . Let  $z'_1, \dots, z'_{c_{x'}}$  be the child nodes of  $x'$ . For  $i \in \{1, \dots, c_{x'}\}$ , if  $z'_i$  is a leaf node  $\text{att}(z'_i) \in \mathcal{N}$ , set  $x$ 's child node  $z$  like  $z = \text{map}(z'_i)$  where  $\text{index}(z) = \text{index}(z'_i)$ ; if  $z'_i$  is a non-leaf node, set  $x$ 's child node  $z$  such that  $z = \text{map}(z'_i)$  where  $\text{index}(z) = \text{index}(z'_i) + n + c - 1$ . This procedure is performed recursively, until each node in  $\mathcal{T}$  is mapped to a corresponding node in  $\mathcal{T}_u$ . Notice that in a  $(d, c)$ -bounded access tree, each non-leaf node  $x$  in  $\mathcal{T}$  has at most  $c$  non-leaf child nodes, and this recursive procedure can be terminable.

### 3.3.2 Security Models

#### Bounded Ciphertext-policy Encryption Model.

**Definition 3.3.2.** A BCPABE scheme includes a tuple of probabilistic polynomial-time algorithms as follows:

- **Setup** $(d, c) \rightarrow (\text{PP}, \text{MK})$ : On input an implicit security parameter  $1^k$  and two system parameters  $(d, c)$ , the setup algorithm **Setup** outputs a public parameter PP and a master key MK.
- **KGen** $(\gamma, \text{MK}) \rightarrow (D)$ : On input an attribute set  $\gamma$  and a master key MK, the key generation algorithm **KGen** outputs a secret key  $D$ .
- **Enc** $(\text{PP}, \mathcal{T}, m) \rightarrow (E)$ : On input the public parameter PP, a  $(d, c)$ -bounded access tree  $\mathcal{T}$ , and a message  $m$ , the encryption algorithm **Enc** outputs a ciphertext

$E$ .

- **Dec**( $D, E$ )  $\rightarrow (m)$ : On input a secret key  $D$  and a ciphertext  $E$ , if the attribute set in  $D$  satisfies the access tree in  $E$ , the decryption algorithm **Dec** decrypts the ciphertext  $E$  and returns a message  $m$ ; otherwise, it outputs " $\perp$ ".

**Selective-Tree Security Model for BCPABE.** This model was first introduced by GJPS. The analogous selective-ID model can be found in [34, 37, 59, 84]. A BCPABE scheme is secure in the selective-tree CPA model if no probabilistic polynomial-time adversary  $\mathcal{A}$  has a non-negligible advantage in winning the following game.

**Init**  $\mathcal{A}$  chooses an access tree  $\mathcal{T}^*$  that he wishes to be challenged upon. The challenger runs **Setup** algorithm and gives  $\mathcal{A}$  the resulting public parameter PP. It keeps the corresponding master key MK.

**Phase 1**  $\mathcal{A}$  issues queries for secret keys related with many attribute sets  $\gamma_j$ , where  $\gamma_j$  does not satisfy the access tree  $\mathcal{T}^*$  for all  $j$ .

**Challenge** Once  $\mathcal{A}$  decides that Phase 1 is over, it outputs two messages of same length,  $m_0$  and  $m_1$ , from the message space. The challenger chooses  $\mu \in \{0, 1\}$  at random and encrypts  $m_\mu$  with  $\mathcal{T}^*$ . Then, the ciphertext  $C^*$  is given to  $\mathcal{A}$ .

**Phase 2** The same as Phase 1.

**Guess**  $\mathcal{A}$  outputs a guess  $\mu' \in \{0, 1\}$  and wins the game if  $\mu' = \mu$ .

We define  $\mathcal{A}$ 's advantage in this game as  $|\Pr[\mu' = \mu] - \frac{1}{2}|$ . The selective-tree CPA model can be extended to handle chosen-ciphertext attacks by allowing for decryption queries in Phases 1 and 2, denoted as selective-tree CCA model.

### 3.3.3 Basic BCPABE Scheme $\text{BCP}_1$

We now proceed with the formal description of our first scheme  $\text{BCP}_1$ .

**Setup:** ( $d, c$ ) This algorithm takes two parameters ( $d, c$ ) as its input. Define a real attribute set  $\mathcal{U} = \{1, \dots, n\}$  and a dummy attribute set  $\mathcal{U}^* = \{n+1, \dots, n+c-1\}$ . Next,



we define a  $(d, c)$ -universal access tree  $\mathcal{T}_u$  as explained in Section 2.4.  $(d, c, \mathcal{U}, \mathcal{U}^*, \mathcal{T}_u)$  are all used in the following algorithms.

Now, the algorithm generates the public parameter for this scheme. For each real attribute  $j \in \mathcal{U}$ , randomly choose a set of  $|\Sigma_{\mathcal{T}_u}|$  numbers  $\{t_{j,x}\}_{x \in \Sigma_{\mathcal{T}_u}}$  from  $Z_p$ . Further, for each dummy attribute  $j \in \mathcal{U}^*$ , randomly choose a set of  $|\Sigma_{\mathcal{T}_u}|$  numbers  $\{t_{j,x}^*\}_{x \in \Sigma_{\mathcal{T}_u}}$  from  $Z_p$ . Finally, randomly choose  $y \in Z_p$ . The public parameter  $\text{PP} = \langle Y = e(g, g)^y, \{T_{j,x} = g^{t_{j,x}}\}_{j \in \mathcal{U}, x \in \Sigma_{\mathcal{T}_u}}, \{T_{j,x}^* = g^{t_{j,x}^*}\}_{j \in \mathcal{U}^*, x \in \Sigma_{\mathcal{T}_u}} \rangle$ . The master key  $\text{MK} = \langle y, \{t_{j,x}\}_{j \in \mathcal{U}, x \in \Sigma_{\mathcal{T}_u}}, \{t_{j,x}^*\}_{j \in \mathcal{U}^*, x \in \Sigma_{\mathcal{T}_u}} \rangle$ .

**KGen**( $\gamma, \text{MK}$ ): This algorithm takes an attribute set  $\gamma \subseteq \mathcal{U}$  and the master key  $\text{MK}$  as its input, then it outputs a secret key  $D$  which can be used for decrypting a ciphertext encrypted under a  $(d, c)$ -bounded access tree  $\mathcal{T}$  iff  $\mathcal{T}(\gamma) = 1$ .

Now, the algorithm generates the secret key. For each user, choose a random polynomial  $q_x$  of degree  $c-1$  for each non-leaf node  $x$  in the  $(d, c)$ -universal access tree. These polynomials are chosen in a top-down manner, satisfying  $q_x(0) = q_{p(x)}(\text{index}(x))$  and  $q_r(0) = y$ . Once the polynomials have been fixed, it outputs the following secret key  $D = \langle \gamma, \{D_{j,x} = g^{\frac{q_x(j)}{t_{j,x}}}\}_{j \in \gamma, x \in \Sigma_{\mathcal{T}_u}}, \{D_{j,x}^* = g^{\frac{q_x(j)}{t_{j,x}^*}}\}_{j \in \mathcal{U}^*, x \in \Sigma_{\mathcal{T}_u}} \rangle$ .

**Enc**( $m, \text{PP}, \mathcal{T}$ ): This algorithm takes a message  $m$ , the public parameter  $\text{PP}$ , and a  $(d, c)$ -bounded access tree  $\mathcal{T}$  as its input.

Now, to encrypt the message  $m$  with the access tree  $\mathcal{T}$ , the algorithm first sets a map from  $\mathcal{T}$  to  $\mathcal{T}_u$  using the method mentioned in Section 2.4. Then, for each non-leaf node  $x \in \mathcal{T}$ , it chooses an arbitrary  $(c - k_x)$ -sized set  $\omega_x$ <sup>4</sup> of dummy child nodes of  $x'$  in  $\mathcal{T}_u$ ,  $x' = \text{map}(x)$ . Following this, let  $f(j, x) = 1$  if the node  $x$  in  $\mathcal{T}$  has a child node associated with real attribute  $j$ ;  $f(j, x) = 0$  otherwise. Then, it chooses a random value  $s \in Z_p$  and outputs the ciphertext  $E = \langle \mathcal{T}, E' = m \cdot Y^s, \{E_{j,x} =$

<sup>4</sup>Without loss of generality,  $w_x = \{n+1, \dots, n+c-k_x\}$ .

$$T_{j,\text{map}(x)}^s\}_{j \in \mathcal{U}, x \in \Sigma_{\mathcal{T}}, f(j,x)=1}, \{E_{j,x}^* = T_{j,\text{map}(x)}^{*s}\}_{j \in \omega_x, x \in \Sigma_{\mathcal{T}}}.$$

**Dec**( $E, D$ ): This algorithm takes a ciphertext  $E$  and a secret key  $D$  as its input. If the attribute set  $\gamma$  associated with  $D$  satisfies the access tree  $\mathcal{T}$  in  $E$ , the algorithm proceeds as follows; otherwise, output  $\perp$ .

A recursive algorithm **DecryptNode**( $E, D, x$ ) takes the ciphertext  $E$ , the secret key  $D$ , and a satisfied non-leaf node  $x$  in  $\mathcal{T}$  as its input and outputs a group element of  $G_T$  or  $\perp$ . For each  $x$ 's child node  $z$ ,

- $z$  is a real node, let  $j = \text{att}(z)$ . Then, we have

$$\begin{aligned} F_{x,j} &= \text{DecryptNode}(E, D, x) \\ &= \begin{cases} e(D_{j,\text{map}(x)}, E_{j,x}) = e(g, g)^{sq_{\text{map}(x)}(j)}, & \text{if } j \in \gamma; \\ \perp, & \text{otherwise.} \end{cases} \end{aligned}$$

- $z$  is a dummy node.  $j = \text{att}(z) \in w_x$ . Then we have:

$$\begin{aligned} F_{x,j} &= \text{DecryptNode}(E, D, x) \\ &= e(D_{j,\text{map}(x)}^*, E_{j,x}^*) = e(g, g)^{sq_{\text{map}(x)}(j)}, \end{aligned}$$

From the above procedure, for each non-leaf node  $x$  in  $\mathcal{T}$ , if  $\mathcal{T}_x(\gamma) = 1$ , we have at least  $k_x + c - k_x = c$  different points  $F_{x,j}$  to compute  $e(g, g)^{sq_{\text{map}(x)}(0)}$  using Lagrange interpolation. By recursively executing such procedure in a bottom-up manner, and finally, it obtains  $E'' = e(g, g)^{sq_r(0)} = e(g, g)^{sy}$ , where  $r$  is the root of  $\mathcal{T}$ . The decryption algorithm outputs  $m = E'/E''$ .

### 3.3.4 Security Proof of $\text{BCP}_1$

**Theorem 3.3.3.** *If the DBDH assumption holds in  $(G, G_T)$ , then scheme  $\text{BCP}_1$  is selective-tree CPA secure in the standard model.*

*Proof.* Suppose a polynomial-time adversary  $\mathcal{A}$  exists that can attack  $\text{BCP}_1$  in the selective-tree CPA model with non-negligible advantage  $\varepsilon$ . We construct a simulator

$\mathcal{S}$  that can distinguish the DBDH tuple from a random tuple with non-negligible advantage  $\frac{\epsilon}{2}$ .

We first let the challenger set the groups  $G$  and  $G_T$  with an efficient bilinear map  $e$  and a generator  $g$ . The challenger flips a fair binary coin  $\nu$ , outside of  $\mathcal{S}$ 's view. If  $\nu = 1$ , the challenger sets  $(g, A, B, C, Z) \in \mathcal{D}_{bdh}$ ; otherwise it sets  $(g, A, B, C, Z) \in \mathcal{D}_{rand}$ .

**Init** The simulator  $\mathcal{S}$  runs  $\mathcal{A}$ .  $\mathcal{A}$  chooses a challenge  $(d, c)$ -bounded access tree  $\mathcal{T}^*$ . The simulator  $\mathcal{S}$  sets  $Y = e(A, B) = e(g, g)^{ab}$ . Then,  $\mathcal{S}$  generates a  $(d, c)$ -universal access tree  $\mathcal{T}_u$  and a map from  $\mathcal{T}^*$  to  $\mathcal{T}_u$ . Randomly choose

$$\{r_{j,x}\}_{j \in \mathcal{U}, x \in \Sigma_{\mathcal{T}_u}}, \{r_{j,x}^*\}_{j \in \mathcal{U}^*, x \in \Sigma_{\mathcal{T}_u}}$$

from  $Z_p$ .

For  $j \in \mathcal{U}, x \in \Sigma_{\mathcal{T}_u}$ ,

$$T_{j,x} = \begin{cases} g^{r_{j,x}}, & \text{if } x = \text{map}(x'), x' \in \Sigma_{\mathcal{T}^*}, f(j, x') = 1; \\ B^{r_{j,x}}, & \text{otherwise.} \end{cases}$$

For  $j \in \mathcal{U}^*, x \in \Sigma_{\mathcal{T}_u}$ ,

$$T_{j,x}^* = \begin{cases} g^{r_{j,x}^*}, & \text{if } j \in w_{x'}, x = \text{map}(x'), x' \in \Sigma_{\mathcal{T}^*}; \\ B^{r_{j,x}^*}, & \text{otherwise.} \end{cases}$$

**Phase 1**  $\mathcal{A}$  adaptively makes query for a secret key related to an attribute set  $\gamma$  such that  $\mathcal{T}^*(\gamma) = 0$ . To generate the secret key,  $\mathcal{S}$  needs to assign a polynomial  $q_x$  for every non-leaf node in  $\mathcal{T}$  and output a piece of secret key according to each non-leaf node.

With an attribute set  $\gamma$  as input, for any node  $x \in \mathcal{T}_u$ , we call a node  $x$ : an unsatisfied node iff there exists an unsatisfied node  $x'$  in  $\mathcal{T}^*$  such that  $\text{map}(x') = x$ ; a satisfied node iff there exists a satisfied node  $x'$  in  $\mathcal{T}^*$  such that  $\text{map}(x') = x$ ; a non-mapped

node iff there exists no node  $x'$  in  $\mathcal{T}^*$  so that  $\text{map}(x') = x$ . We define the following three procedures: PolyUnsat, PolySat, and PolyNotCare.

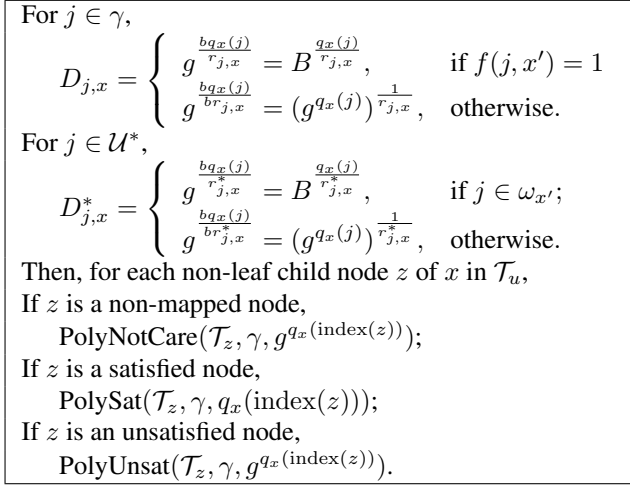


Figure 3.6: PolyUnsat( $\mathcal{T}_x, \gamma, g^{\lambda_x}$ )

PolyUnsat( $\mathcal{T}_x, \gamma, g^{\lambda_x}$ ) for an unsatisfied node  $x \in \Sigma_{\mathcal{T}}$  is defined as follows:

This procedure generates a polynomial  $q_x$  for an unsatisfied node  $x$ . We have an unsatisfied node  $x'$  such that  $\text{map}(x') = x$ .  $\gamma$  does not satisfy this access tree  $\mathcal{T}_{x'}$ , denoted as  $\mathcal{T}_{x'}^*(\gamma) = 0$ , where  $\mathcal{T}_{x'}^*$  is a subtree of  $\mathcal{T}^*$ .  $\lambda_x$  is an integer from  $Z_p$ . The unsatisfied root node  $x'$  has at most  $k_{x'} - 1$  satisfied child node. Thus, it could implicitly sets  $q_x(0) = \lambda_x$ , and chooses  $c - 1$  other points at random to fix  $q_x$  completely, including

- $c - k_{x'}$  points as  $q_x(j)$  for the dummy nodes of  $x$ , where  $j \in w_{x'}$ ;
- At most  $k_{x'} - 1$  points as  $q_x(\text{index}(z'))$  if  $z'$  is a satisfied leaf child node of  $x'$  or  $q_x(\text{index}(z') + n + c - 1)$  where  $z'$  is a satisfied non-leaf child node of  $x'$ .

It executes the following steps in Figure 3.6.

PolySat( $\mathcal{T}_x, \gamma, \lambda_x$ ) for a satisfied node  $x \in \Sigma_{\mathcal{T}}$  is defined as follows:

This procedure generates a polynomial  $q_x$  for a satisfied node  $x$ . We have a satisfied node  $x'$  such that  $\text{map}(x') = x$ .  $\lambda_x$  is an integer from  $Z_p$ . It sets  $q_x(0) = \lambda_x$ , and chooses  $c - 1$  other points at random to completely fix  $q_x$ . Thus, for each  $j \in \mathcal{U} \cup \mathcal{U}^* \cup \{n + c, \dots, n + 2c - 1\}$ , we can obtain  $q_x(j)$ .

It executes the following steps in Figure 3.7.

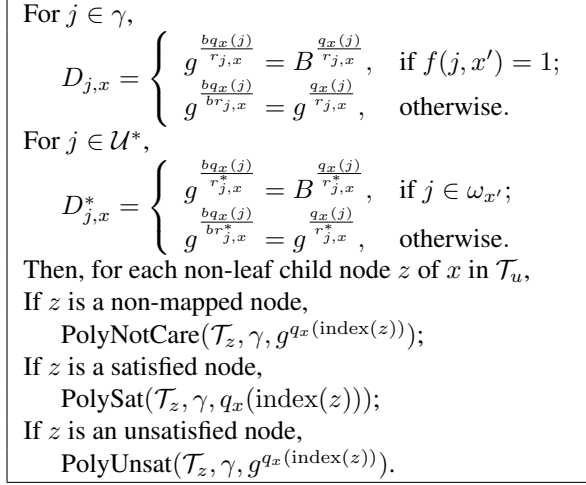


Figure 3.7:  $\text{PolySat}(\mathcal{T}_x, \gamma, \lambda_x)$

$\text{PolyNotCare}(\mathcal{T}_x, \gamma, g^{\lambda_x})$  for a non-mapped node  $x \in \Sigma_{\mathcal{T}}$  is defined as follows:

This procedure generates a polynomial  $q_x$  for a non-mapped node  $x$ . It implicitly sets  $q_x(0) = \lambda_x$ , and chooses  $c - 1$  other points at random to fix  $q_x$  implicitly. Thus, for  $j \in \mathcal{U} \cup \mathcal{U}^* \cup \{n + c, \dots, n + 2c - 1\}$ , we could obtain  $g^{q_x(j)}$ . It outputs the following secret keys:

$$\{D_{j,x} = g^{\frac{bq_x(j)}{br_{j,x}}} = (g^{q_x(j)})^{\frac{1}{r_{j,x}}}\}_{j \in \gamma}, \{D_{j,x}^* = g^{\frac{bq_x(j)}{br_{j,x}^*}} = (g^{q_x(j)})^{\frac{1}{r_{j,x}^*}}\}_{j \in \mathcal{U}^*}$$

Then, it calls  $\text{PolyNotCare}(\mathcal{T}_z, \gamma, g^{q_x(\text{index}(z))})$  for each non-leaf child node  $z$  of  $x$  in  $\mathcal{T}_u$ .

To give a secret key for an attribute set  $\gamma$ ,  $\mathcal{S}$  first runs  $\text{PolyUnsat}(\mathcal{T}_r = \mathcal{T}, \gamma, A)$ . Notice that we implicitly set  $y = ab$  by  $Y = e(A, B) = e(g, g)^y$ . The secret key

corresponding to each non-leaf node is recursively given by the above three procedures. Finally, it outputs

$$D = \langle \gamma, \{D_{j,x}\}_{j \in \gamma, x \in \Sigma_{\mathcal{T}_u}}, \{D_{j,x}^*\}_{j \in \mathcal{U}^*, x \in \Sigma_{\mathcal{T}_u}} \rangle$$

Therefore,  $\mathcal{S}$  can answer each secret key query with an attribute set  $\gamma$ , where  $\mathcal{T}^*(\gamma) = 0$ . The distribution of these secret keys are identical to those in the real environment.

**Challenge** The adversary  $\mathcal{A}$  will submit two challenge messages  $m_0$  and  $m_1$  to  $\mathcal{S}$ . Then,  $\mathcal{S}$  chooses  $\mu \in \{0, 1\}$  at random, and returns an encryption of  $m_\mu$  under the challenge access tree  $\mathcal{T}^*$ . The challenge ciphertext  $E$  is formed as:

$$\begin{aligned} \langle \mathcal{T}^*, E' = m_\mu \cdot Z, \{E_{j,x} = C^{r_{j,\text{map}(x)}}\}_{j \in \mathcal{U}, x \in \Sigma_{\mathcal{T}^*}, f(j,x)=1}, \\ \{E_{j,x}^* = C^{r_{j,\text{map}(x)}^*}\}_{j \in \omega_x, x \in \Sigma_{\mathcal{T}^*}} \rangle \end{aligned}$$

If  $(g, A, B, C, Z) \in \mathcal{D}_{bdh}$  and we let  $s = c$ , we have  $Y^s = e(g, g)^{abc}$  and

$$E_{j,x} = C^{r_{j,\text{map}(x)}} = (g^{r_{j,\text{map}(x)}})^s, E_{j,x}^* = C^{r_{j,\text{map}(x)}^*} = (g^{r_{j,\text{map}(x)}^*})^s.$$

Therefore, the ciphertext is a valid random encryption of message  $m_\mu$ .

Otherwise, if  $(g, A, B, C, Z) \in \mathcal{D}_{rand}$ , we have  $E' = m_\mu \cdot Z$ . Since  $Z$  is randomly chosen from  $G_T$ ,  $E'$  will be a random element of  $G_T$  from the adversary's view and the ciphertext doesn't contain any information about  $m_\mu$ .

**Phase 2** The simulator  $\mathcal{S}$  acts exactly as it did in **Phase 1**.

**Guess**  $\mathcal{S}$  outputs  $\nu' = 1$  to indicate that it was given a tuple from  $\mathcal{D}_{bdh}$  if  $\mathcal{A}$  gives a

correct guess  $\mu' = \mu$ ; otherwise output  $\nu' = 0$  to indicate that it was given a tuple from  $\mathcal{D}_{rand}$ .

Let us compute the success probability of  $\mathcal{S}$ :

In the case of  $\nu = 0$ , the adversary gains no information about  $\mu$ . Therefore, we have  $\Pr[\mu \neq \mu' | \nu = 0] = \frac{1}{2}$ . Since the simulator guesses  $\nu' = 0$  when  $\mu \neq \mu'$ , we have  $\Pr[\nu' = \nu | \nu = 0] = \Pr[\nu' = 0 | \nu = 0] = \frac{1}{2}$ .

In the case of  $\nu = 1$ , the adversary gets a valid ciphertext of  $m_\mu$ . By definition, the adversary has probability  $\varepsilon$  to guess the correct  $\mu'$ , and thus  $\Pr[\mu = \mu' | \nu = 1] = \frac{1}{2} + \varepsilon$ . Since the simulator guesses  $\nu' = 1$  when  $\mu = \mu'$ , we have  $\Pr[\nu' = \nu | \nu = 1] = \Pr[\nu' = 1 | \nu = 1] = \frac{1}{2} + \varepsilon$ .

The overall advantage of the simulator to output a correct  $\nu' = \nu$  is  $\Pr[\nu = \nu'] - \frac{1}{2} = \Pr[\nu = \nu', \nu = 0] + \Pr[\nu = \nu', \nu = 1] - \frac{1}{2} = \frac{1}{2} \cdot \frac{1}{2} + \frac{1}{2} \cdot (\frac{1}{2} + \varepsilon) - \frac{1}{2} = \frac{\varepsilon}{2}$   $\square$

### 3.3.5 Extended BCPABE Scheme $\text{BCP}_2$

Now, by using one-time signature technique, we present an extended scheme  $\text{BCP}_2$  achieving chosen ciphertext security.

The selective-tree CCA model was introduced in Section 3.3.2 and the similar security model can be found in [84].

We assume that existing BCPABE scheme  $\text{BCP}_1$  is secure in the selective-tree CPA model as presented in Section 3.3.2, including four algorithms

$$(\text{BCP}_1.\mathbf{Setup}, \text{BCP}_1.\mathbf{KGen}, \text{BCP}_1.\mathbf{Enc}, \text{BCP}_1.\mathbf{Dec})$$

and a secure one-time signature  $\text{ots}$ , including three algorithms

$$(\text{ots}.\mathbf{KGen}, \text{ots}.\mathbf{Sig}, \text{ots}.\mathbf{Ver}).$$

Assume that the verification key  $vk$  from  $\text{ots}$  is a bit string of length  $l$ , and we write  $vk_i$  for the  $i$ -th bit in  $vk$ . Let  $\mathcal{L}$  denote  $\{1, 2, \dots, l\}$ .  $\text{BCP}_2$  is constructed based on  $\text{BCP}_1$  and  $\text{ots}$  including the following algorithms:

**Setup**( $d, c$ ): This algorithm takes two system parameters  $(d, c)$  as its input. Then, it calls  $\text{BCP}_1.\text{Setup}(d, c)$  to generate  $\text{BCP}_1$ 's public parameter  $\text{PP}_1$  and master key  $\text{MK}_1$ . In addition, it randomly chooses a set  $\{t'_i\}_{i \in \{1, \dots, 2l\}}$  from  $Z_p$  and defines  $T'_i = g^{t'_i}$ . Now, it outputs the public parameter  $\text{PP} = \langle \text{PP}_1, \{T'_i\}_{i \in \{1, \dots, 2l\}} \rangle$  and keeps the master key  $\text{MK} = \langle \text{MK}_1, \{t'_i\}_{i \in \{1, \dots, 2l\}} \rangle$ .

**KGen**( $\gamma, \text{MK}$ ): This algorithm takes an attribute set  $\gamma$  and the master key  $\text{MK}$  as its input. Then, it randomly chooses  $r'$  from  $Z_p$  and calls  $\text{BCP}_1.\text{KGen}(\gamma, \text{MK}_1)$  to generate a user's secret key  $D_1$  by using  $r'$  instead of  $y$  in  $\text{MK}_1$  (i.e.,  $q_r(0) = r'$ ). For every  $i \in \mathcal{L}$ , let  $D_{i,0} = g^{\frac{r_i}{t'_i}}$  and  $D_{i,1} = g^{\frac{r_i}{t'_{l+i}}}$ , where  $\{r_i\}_{i \in \mathcal{L}}$  are randomly chosen from  $Z_p$ . Define  $r = r' + \sum_{i \in \mathcal{L}} r_i$  and let  $\hat{D} = g^{y-r}$ . Finally, it outputs  $D = \langle D_1, \{D_{i,0}, D_{i,1}\}_{i \in \mathcal{L}}, \hat{D} \rangle$ .

**Enc**( $m, \text{PP}, \mathcal{T}$ ): This algorithm takes a message  $m$ , the public parameter  $\text{PP}$ , and a  $(d, c)$ -bounded access tree  $\mathcal{T}$  as its input.

It first calls  $\text{BCP}_1.\text{Enc}(m, \text{PP}_1, \mathcal{T})$  and obtains a partial ciphertext  $E_1$ . Then, a key pair  $\langle sk, vk \rangle$  is obtained by running  $\text{ots.KGen}$ . For each  $i \in \mathcal{L}$ , it sets  $E'_i = T'_i{}^s$  if  $vk_i = 0$ ;  $E'_i = T'_{l+i}{}^s$  otherwise. Let  $\hat{E} = g^s$ .<sup>5</sup> It runs  $\text{ots.Sig}$  with input  $(sk, \langle E_1, \{E'_i\}_{i \in \mathcal{L}}, \hat{E} \rangle)$  and obtains  $\sigma$ .

The output ciphertext  $E = \langle E_1, \{E'_i\}_{i \in \mathcal{L}}, \hat{E}, \sigma, vk \rangle$ .

**Dec**( $D, E$ ): This algorithm takes a secret key  $D$  and a ciphertext  $E$  as its input.

It first checks whether  $\sigma$  is a valid signature on message  $\langle E_1, \{E'_i\}_{i \in \mathcal{L}}, \hat{E} \rangle$  using  $vk$ .

---

<sup>5</sup>Here,  $s$  is consistent with the random value in  $E_1$



If valid, it proceeds as follows; otherwise, output  $\perp$ .

It extracts  $D_1$  and  $E_1$  from tuple  $(D, E)$  and consequently decrypts  $e(g, g)^{sr'}$  according to  $\text{BCP}_1$ . **Dec.**

For each  $i \in \mathcal{L}$ , it computes

$$e(g, g)^{sr_i} = \begin{cases} e(D_{i,0}, T_i'^s) = e(g^{\frac{r_i}{t_i}}, g^{t_i'^s}), & \text{if } vk_i = 0; \\ e(D_{i,1}, T_{l+i}'^s) = e(g^{\frac{r_i}{t_{l+i}'}}, g^{t_{l+i}'^s}), & \text{if } vk_i = 1. \end{cases}$$

Finally, it computes  $m = \frac{E'}{e(\hat{E}, \hat{D}) \cdot e(g, g)^{sr'} \cdot \prod_{i \in \mathcal{L}} e(g, g)^{sr_i}}$ .

Compared with  $\text{BCP}_1$ ,  $\text{BCP}_2$ 's ciphertext is augmented with  $l$  elements, while public parameter and secret key are both augmented with  $2l$  elements. Another method for extending  $\text{BCP}_1$  with CCA security level has been mentioned [129]. It treats each verification key as an attribute. However, it has shorter additional size of ciphertext (1 element) but larger additional size of public parameter and secret key ( $2^l$  elements).

### 3.3.6 Security Proof of $\text{BCP}_2$

The selective-tree CCA model is introduced in Section 3.3.2. We prove the security of  $\text{BCP}_2$  based on the strong existentially unforgeable assumption of  $\text{ots}$  and the DBDH assumption.

**Theorem 3.3.4.** *Suppose  $\text{ots}$  is a  $\varepsilon_{\text{ots}}$ -secure one-time signature scheme. If the DBDH assumption holds in  $(G, G_T)$ , then scheme  $\text{BCP}_2$  is a selective-tree CCA secure in the standard model.*

*Proof.* Suppose there is a polynomial-time adversary  $\mathcal{A}$  who can attack  $\text{BCP}_2$  in the selective-tree CCA model with non-negligible advantage  $\varepsilon$ . We construct a simulator  $S$  who can distinguish the DBDH tuple from a random tuple with non-negligible advantage  $\frac{\varepsilon}{2} - \varepsilon_{\text{ots}}$ .

We first let the challenger set the groups  $G$  and  $G_T$  with an efficient bilinear map  $e$

and a generator  $g$ . The challenger flips a fair binary coin  $\nu$ , outside of  $\mathcal{S}$ 's view. If  $\nu = 1$ , the challenger sets  $(g, A, B, C, Z) \in \mathcal{D}_{bdh}$ ; otherwise sets  $(g, A, B, C, Z) \in \mathcal{D}_{rand}$ .

**Init** The simulator  $\mathcal{S}$  runs  $\mathcal{A}$ .  $\mathcal{A}$  chooses a  $(d, c)$ -bounded access tree  $\mathcal{T}^*$  it wishes to be challenged upon.  $\mathcal{S}$  runs  $\text{OTS.KGen}$  to obtain  $\langle sk^*, vk^* \rangle$ .  $\mathcal{S}$  sets  $Y = e(g, g)^{ab} = e(A, B)$ .  $\mathcal{S}$  defines a  $(d, c)$ -universal access tree  $\mathcal{T}$  and a map from  $\mathcal{T}^*$  to  $\mathcal{T}$ . Then, it generates  $\text{PP}_1$  as **Init** step in Section 4.2.

For  $i \in \mathcal{L}$ , randomly choose  $\eta_i, \theta_i \in Z_p$  and implicitly set

if  $vk_i^* = 0$ ,

$$t'_i = \eta_i, T'_i = g^{\eta_i} \text{ and } t'_{l+i} = b\theta_{l+i}, T'_{l+i} = B^{\theta_{l+i}};$$

if  $vk_i^* = 1$ ,

$$t'_i = b\eta_i, T'_i = B^{\eta_i} \text{ and } t'_{l+i} = \theta_{l+i}, T'_{l+i} = g^{\theta_{l+i}}.$$

The algorithm outputs public parameter

$$\text{PP} = \langle \text{PP}_1, \{T'_i\}_{i \in \{1, \dots, 2l\}} \rangle$$

**Phase 1**  $\mathcal{A}$  is allowed to make secret key queries and decryption queries:

- **Secret Key Query.**  $\mathcal{A}$  submits an attribute set  $\gamma$  such that  $\mathcal{T}^*(\gamma) = 0$ .  $\mathcal{S}$  randomly chooses  $r'', r'_i \in Z_p$  for  $i \in \mathcal{L}$  and implicitly sets:  $r' = ab + br'', r_i = br'_i$ .

According to the Phase 1 in Section 4, it calls  $\text{PolyUnsat}(\mathcal{T}_u, \gamma, A \cdot g^{r''})$  and obtains  $D_1$ . Then, it computes

$$\hat{D} = g^{y-r} = g^{ab-ab-br''-\sum_{i \in \mathcal{L}} br'_i} = \frac{1}{B^{r''+\sum_{i \in \mathcal{L}} r'_i}}$$

and for  $i \in \mathcal{L}$ , if  $vk_i^* = 0$ ,

$$D_{i,0} = B^{\frac{r'_i}{\eta_i}} \text{ and } D_{i,1} = g^{\frac{r'_i}{\theta_{t+i}}};$$

if  $vk_i^* = 1$

$$D_{i,0} = g^{\frac{r'_i}{\eta_i}} \text{ and } D_{i,1} = B^{\frac{r'_i}{\theta_{t+i}}}.$$

Finally, it outputs the secret key

$$D = \langle D_1, \{D_{i,0}, D_{i,1}\}_{i \in \mathcal{L}}, \hat{D} \rangle.$$

- **Decryption Query.**  $\mathcal{A}$  submits a ciphertext  $E = \langle E_1, \{E'_i\}_{i \in \mathcal{L}}, \hat{E}, \sigma, vk \rangle$  related with  $\mathcal{T}$ .  $\mathcal{S}$  checks the signature  $\sigma$  using  $vk$ . If  $\sigma$  is invalid,  $\mathcal{S}$  outputs  $\perp$ ; otherwise  $\mathcal{S}$  checks if  $vk = vk^*$ . If so, we call it a **forge** event and  $\mathcal{S}$  outputs  $\nu' = 0$  to indicate that it was given a tuple from  $\mathcal{D}_{rand}$ . Now, the only case is  $vk \neq vk^*$ . In such a case,  $\mathcal{S}$  defines an attribute set  $\gamma$  such that  $\mathcal{T}(\gamma) = 1$ . If  $\mathcal{T}^*(\gamma) = 0$ , generate a secret key related with  $\gamma$  from secret key query and use it to decrypt  $E$ ; otherwise  $\mathcal{T}^*(\gamma) = 1$ . Without loss of generality, assume  $vk_j = 1, vk_j^* = 0$ .  $\mathcal{S}$  generates a partial secret key for decrypting  $E$  as follows:

- $\mathcal{S}$  randomly chooses  $r'', r'_i \in \mathbb{Z}_p$  for  $i \in \mathcal{L}$  and implicitly sets:  $r' = br'', r_i = br'_i$  for  $i \neq j$  and  $r_j = ab + br'_j$ .
- According to the Phase 1 in Section 4.2, it calls  $\text{PolySat}(\mathcal{T}, \gamma, r'')$  and obtains  $D_1$  related with  $\gamma$ . Now, it can use  $D_1$  to decrypt  $E_1$  and obtains  $e(g, g)^{sr'}$ .
- For  $i \in \mathcal{L}$  and  $i \neq j$ , if  $vk_i^* = 0$ ,

$$D_{i,0} = B^{\frac{r'_i}{\eta_i}} \text{ and } D_{i,1} = g^{\frac{r'_i}{\theta_{t+i}}};$$

if  $vk_i^* = 1$ ,

$$D_{i,0} = g^{\frac{r'_i}{\eta_i}} \text{ and } D_{i,1} = B^{\frac{r'_i}{\theta_{l+i}}}.$$

- For  $i = j$ , it can only generate  $D_{j,1} = g^{\frac{r_j}{\eta_{l+j}}} = g^{\frac{ab+br'_j}{b\theta_{l+j}}} = A^{\frac{1}{\theta_{l+j}}} \cdot g^{\frac{r'_j}{\theta_{l+j}}}$ .

This is sufficient for our decryption since  $vk_j = 1$  and  $T_{l+j}^{rs}$  is involved in ciphertext  $E$ . Note that  $e(D_{j,1}, T_{l+j}^{rs}) = e(g, g)^{sr_j}$ .

- $\mathcal{S}$  then computes  $\hat{D} = g^{y-r} = g^{ab-br''-ab-\sum_{i \in \mathcal{L}} br'_i} = \frac{1}{B^{r''+\sum_{i \in \mathcal{L}} r'_i}}$ .
- Finally,  $\mathcal{S}$  outputs  $m = \frac{E'}{e(\hat{E}, \hat{D}) \cdot e(g, g)^{sr'} \cdot \prod_{i \in \mathcal{L}} e(g, g)^{sr_i}}$ .

**Challenge** The adversary  $\mathcal{A}$  will submit two challenge messages  $m_0$  and  $m_1$  to  $\mathcal{S}$ . Then,  $\mathcal{S}$  chooses  $\mu \in \{0, 1\}$  at random, and returns an encryption of  $m_\mu$  under the challenge access tree  $\mathcal{T}^*$  as follows:

- It first generates an encryption  $E_1^*$  according to **Challenge** in Section 4.2.
- It generates a signature

$$\sigma^* = \text{ots.Sig}(sk^*, \langle E_1^*, \{C^{\eta_i}\}_{i \in \mathcal{L}, vk_i^*=0}, \{C^{\theta_{l+i}}\}_{i \in \mathcal{L}, vk_i^*=1}, C \rangle).$$

- It outputs the challenge ciphertext

$$E^* = \langle E_1^*, \{C^{\eta_i}\}_{i \in \mathcal{L}, vk_i^*=0}, \{C^{\theta_{l+i}}\}_{i \in \mathcal{L}, vk_i^*=1}, C, \sigma^*, vk^* \rangle.$$

Let  $s = c$ . If  $(g, A, B, C, Z) \in \mathcal{D}_{bdh}$ , we have

$$Y^s = e(g, g)^{abc}, E'_i = C^{\eta_i} = (g^s)^{\eta_i} = T_i^{rs}, E'_i = C^{\theta_{l+i}} = (g^s)^{\theta_{l+i}} = T_{l+i}^{rs}.$$

Therefore, the ciphertext is a valid random encryption of message  $m_\mu$ .

Otherwise, if  $(g, A, B, C, Z) \in \mathcal{D}_{rand}$ , we have  $E' = m_\mu \cdot Z$ . Since  $Z$  is randomly chosen from  $G_T$ ,  $E'$  will be a random element of  $G_T$  from the adversary's view. Thus

the ciphertext contains no information about  $m_\mu$ .

**Phase 2** The simulator  $\mathcal{S}$  acts exactly as it did in **Phase 1**.

**Guess**  $\mathcal{S}$  outputs  $\nu' = 1$  to indicate that it was given a tuple from  $\mathcal{D}_{bdh}$  if  $\mathcal{A}$  gives a correct guess  $\mu' = \mu$ ; otherwise  $\nu' = 0$  to indicate that it was given a tuple from  $\mathcal{D}_{rand}$ .

Let us compute the success probability of  $\mathcal{S}$ :

In the case of  $\nu = 0$ , the adversary doesn't gain any information about  $\mu$ . Therefore, we have  $\Pr[\mu \neq \mu' | \nu = 0] = \frac{1}{2}$ . Since the simulator guesses  $\nu' = 0$  when  $\mu \neq \mu'$  (no **forge**), we have

$$\begin{aligned} \Pr[\nu' = \nu | \nu = 0] &= \Pr[\nu' = 0 | \nu = 0] = \Pr[\mu \neq \mu', \neg \text{forge} | \nu = 0] + \Pr[\text{forge} | \nu = 0] \\ &= \Pr[\mu \neq \mu' | \nu = 0] - \Pr[\mu \neq \mu', \text{forge} | \nu = 0] + \Pr[\text{forge} | \nu = 0] \\ &\geq \frac{1}{2} - \Pr[\text{forge} | \nu = 0] = \frac{1}{2} - \varepsilon_{ots}. \end{aligned}$$

In the case of  $\nu = 1$ , the adversary gets a valid ciphertext of  $m_\mu$ . By definition, the adversary has probability  $\varepsilon$  to guess the correct  $\mu'$ , and thus  $\Pr[\mu = \mu' | \nu = 1] = \frac{1}{2} + \varepsilon$ . Since the simulator guesses  $\nu' = 1$  when  $(\mu = \mu', \text{no } \text{forge})$ , we have

$$\begin{aligned} \Pr[\nu' = \nu | \nu = 1] &= \Pr[\nu' = 1 | \nu = 1] = \Pr[\mu = \mu', \neg \text{forge} | \nu = 1] \\ &= \Pr[\mu = \mu' | \nu = 1] - \Pr[\mu = \mu', \text{forge} | \nu = 1] \\ &\geq \frac{1}{2} + \varepsilon - \Pr[\text{forge} | \nu = 1] = \frac{1}{2} + \varepsilon - \varepsilon_{ots}. \end{aligned}$$

The overall advantage of the simulator to output a correct  $\nu' = \nu$  is

$$\begin{aligned} \Pr[\nu = \nu'] - \frac{1}{2} &= \Pr[\nu = \nu', \nu = 0] + \Pr[\nu = \nu', \nu = 1] - \frac{1}{2} \\ &\geq \frac{1}{2} \cdot \left(\frac{1}{2} - \varepsilon_{ots}\right) + \frac{1}{2} \cdot \left(\frac{1}{2} + \varepsilon - \varepsilon_{ots}\right) - \frac{1}{2} = \frac{\varepsilon}{2} - \varepsilon_{ots}. \end{aligned}$$

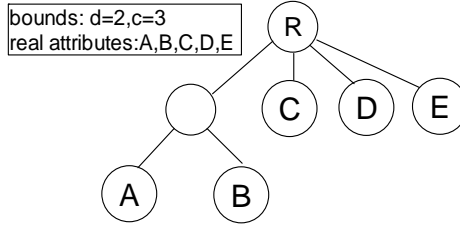


Figure 3.8: An example access tree accepted by  $BCP_1$  but not accepted by GJPS

□

### 3.3.7 Comparisons

**Access Policy:** In this section, we compare the expressive capability of access tree of  $BCP_1$  with that of GJPS scheme bounded by the same parameter  $(d, c)$ . Actually, according to the definition of  $BCP_1$  on the  $(d, c)$ -bounded access tree, each non-leaf node has a threshold value at the most  $c$  and no more than  $c$  non-leaf nodes share a single unique parent. Thus, one difference of the restriction on access trees between GJPS and  $BCP_1$  is if a non-leaf node  $x$  has a non-leaf child node, in GJPS the total number of  $x$ 's child node must be no more than  $c$ , while in  $BCP_1$  the total number of  $x$ 's non-leaf child node must be no more than  $c$ . Therefore, our scheme accommodates more possible access policies chosen by the sender under the same pre-set bounds. (An example is shown in Figure 3.8)

**Efficiency and Parameter Size:** Now, we discuss the comparisons on the computational cost of each algorithm and the sizes of parameters between GJPS and  $BCP_1$  in Table 3.1, both of which are proved secure in the selective-tree CPA model. We assume that both schemes are initialized with same system parameters  $(d, c)$  and consider an encryption under a  $(d, c)$ -bounded access tree  $\mathcal{T}$  (this access tree must be chosen suitable for both schemes since there's a difference between the acceptable

access trees of two schemes).  $\mathcal{T}_n$  is  $\mathcal{T}$ 's normal form. The secret key is associated with an attribute set  $\gamma$  such that  $\mathcal{T}(\gamma) = 1$  and  $|\gamma| = x$ .  $\hat{\mathcal{T}}$  and  $\hat{\mathcal{T}}_n$  are the  $\gamma$ -satisfied non-redundant trees of  $\mathcal{T}$  and  $\mathcal{T}_n$  with minimum non-leaf nodes, respectively.  $|\mathcal{U}| = n, |\mathcal{U}^*| = c - 1, |\Sigma_{\mathcal{T}_u}| = \frac{c^d - 1}{c - 1}$ . Here, **TExp** represents the cost of one modular exponentiation, **TPair** represents the cost of one bilinear pairing computing.  $T_{S1}, T_{K1}, T_{E1}$ , and  $T_{D1}$  represent the computational cost of **Setup**, **KGen**, **Enc** and **Dec** algorithms in GJPS.  $L_{P1}, L_{S1}$ , and  $L_{C1}$  represent the size of public parameter, secret key and ciphertext in GJPS. The mark with “2” indicates the counterpart of  $\text{BCP}_1$ .

Mark	times	Computational cost	
$T_{S1}$	1	$\text{TPair} + (nc^{\frac{d-1}{c-1}} + c^d) \cdot \text{TExp}$	✓
$T_{S2}$	1	$\text{TPair} + (n \times \frac{c^d - 1}{c - 1} + c^d) \cdot \text{TExp}$	
$T_{K1}$	1/user	$(xc^{\frac{d-1}{c-1}} + c^d - 1) \cdot \text{TExp}$	✓
$T_{K2}$	1/user	$(x \times \frac{c^d - 1}{c - 1} + c^d - 1) \cdot \text{TExp}$	
$T_{E1}$	many	$(1 +  \Theta_{\mathcal{T}_n}  + \sum_{x \in \Sigma_{\mathcal{T}_n}} (c - k_x)) \cdot \text{TExp}$	
$T_{E2}$	many	$(1 +  \Theta_{\mathcal{T}}  + \sum_{x \in \Sigma_{\mathcal{T}}} (c - k_x)) \cdot \text{TExp}$	✓
$T_{D1}$	many	$( \Sigma_{\hat{\mathcal{T}}_n}  \times c -  \Sigma_{\hat{\mathcal{T}}_n}  + 1) \cdot \text{TPair} +  \Sigma_{\hat{\mathcal{T}}_n}  \times c \cdot \text{TExp}$	
$T_{D2}$	many	$( \Sigma_{\hat{\mathcal{T}}} \times c -  \Sigma_{\hat{\mathcal{T}}} + 1) \cdot \text{TPair} +  \Sigma_{\hat{\mathcal{T}}} \times c \cdot \text{TExp}$	✓
	number	Size	
$L_{P1}$	1	$ G_T  + (nc^{\frac{d-1}{c-1}} + c^d - 1) \cdot  G $	✓
$L_{P2}$	1	$ G_T  + (n \times \frac{c^d - 1}{c - 1} + c^d - 1) \cdot  G $	
$L_{S1}$	1/user	$(xc^{\frac{d-1}{c-1}} + c^d - 1) \cdot  G $	✓
$L_{S2}$	1/user	$(x \times \frac{c^d - 1}{c - 1} + c^d - 1) \cdot  G $	
$L_{C1}$	many	$ G_T  + ( \Theta_{\mathcal{T}_n}  + \sum_{x \in \Sigma_{\mathcal{T}_n}} (c - k_x)) \cdot  G $	
$L_{C2}$	many	$ G_T  + ( \Theta_{\mathcal{T}}  + \sum_{x \in \Sigma_{\mathcal{T}}} (c - k_x)) \cdot  G $	✓

Table 3.1: Comparisons between the scheme in GJPS and  $\text{BCP}_1$

Generally,  $c$  is set to be no less than 2, then

$$x \times \frac{c^d - 1}{c - 1} + c^d - 1 < 2 \times (xc^{\frac{d-1}{c-1}} + c^d - 1).$$

This is because of the following deduction:

$$2 \leq c \implies 2c^{d-1} - 1 < c^d \implies \frac{c^d - 1}{c - 1} < 2c^{d-1}.$$

Therefore, we obtain that

$$T_{S2} < 2T_{S1}, T_{K2} < 2T_{K1}, L_{P2} < 2L_{P1}, L_{S2} < 2L_{S1}.$$

### 3.4 Multi-Authority Encryption Schemes

When Sahai and Waters [241] introduced the notion of ABE, they also presented the following open problem: Is it possible to construct an ABE scheme in which many different authorities operate simultaneously, each handing out secret keys for a different set of attributes? This is an interesting and practical problem. In the aforementioned ABE systems, all attributes are managed by a single authority. In some applications however, this may not be desirable. For example, Alice encrypts a message with access policy (“UNIV.X.COMPUTER SCIENCE” **AND** “UNIV.X.ALUMNI” **AND** “COMP.Y.ENGINEER”) so that only receivers who are the computer science alumni of University X and currently working as an engineer for Company Y, can decrypt. The authority UNIV.X Registry may only manage attributes for the students, staff, and alumni of University X, while COMP.Y Registry may be the authority handling its employees’ attributes. A single-authority ABE may not be appropriate in this scenario. Another problem of single-authority ABE is the so-called *Key Escrow* problem. In a single-authority ABE system, as the single authority is responsible for issuing private keys for all attributes, it can decrypt any ciphertexts in the system, so that the authority must be fully trusted. Some multi-authority ABE systems [78, 79, 184, 198, 201, 217] have been proposed to achieve better expressiveness, efficiency, and security in the setting of multi-authority. The problems of privacy and key-escrow are also considered simultaneously in those papers. The Table 3.2 shows the properties of these systems in terms of security, expressiveness, and additional properties such as protecting privacy and preventing decryption by individual authority.

Chase [78] proposed the first multi-authority ABE system where there are one CA (Central Authority) and multiple AAs (Attribute Authorities). The CA issues identity-related keys to users and the AAs manage attributes and issue attribute-related keys. A user’s keys from different AAs are linked together by the user’s *global identifier*. The expressiveness of the system is limited and only “AND” policy between the AAs is sup-



Table 3.2: Existing Multi-Authority ABE Systems

	Adaptively Secure	Standard Model	Prevent Decryption by Individual Authorities	KP/CP	Expressiveness
[78]	×	✓	×	KP	Limited <sup>1</sup>
[198]	×	✓	✓	KP	Limited <sup>1</sup>
[79] <sup>4</sup>	×	✓	✓	KP	Limited <sup>1</sup>
[217]	×	✓	×	CP	Expressive <sup>2</sup>
[184]	✓	×	Partially <sup>3</sup>	CP	Expressive <sup>2</sup>
[201]	✓	✓	✓	CP	Expressive <sup>2</sup>

<sup>1</sup> [78, 79, 198] are KPABE systems. In these systems, the policy of each key is defined by some sub-policies, where each sub-policy corresponds to an authority, and only when all sub-policies of a policy are satisfied, the policy is satisfied. i.e., the policy supports *only* **AND** relation between authorities.

<sup>2</sup> [184, 201, 217] are CPABE systems, and the encryptor can encrypt messages with any monotone access structures defined over the *whole attribute universe*.

<sup>3</sup> In [78, 217], there is an authority, called Central Authority, can decrypt *all* ciphertexts. Although no individual authority in [184] can decrypt *all* ciphertexts in the system, each authority can individually decrypt the ciphertexts whose policies are satisfied by the attributes managed by the authority.

<sup>4</sup> The system in [79] can protect user privacy from the authorities.

ported. Also, the CA can decrypt all ciphertexts. We [198] removed the CA using a threshold technique where the set of authorities is fixed in advance and all authorities must interact during the system setup. The system cannot defend against collusion attack by  $m$  or more users where  $m$  is a system parameter chosen at setup. Chase and Chow [79] also removed the central authority using a distributed PRF (pseudo random function) technique. However, the expressiveness is as limited as the original Chase's system, and their technique does not seem applicable to CP-ABE. While [78, 79, 198] focused on KP-ABE, Müller, Katzenbeisser, and Eckert [217] proposed the first multi-authority CP-ABE system where there are one CA and multiple AAs. The AAs operate independently from each other and therefore the scheme is flexible and practical. However, the CA in the system can still decrypt all ciphertexts.

Lewko and Waters [184] proposed the first adaptively secure multi-authority CPABE system. The system is expressive, supporting any monotone access structures. There is no central authority and each authority in the system operates independently.

However, the system is proven secure in the random oracle model, and each authority can still independently decrypt ciphertexts, if the attributes, managed by the authority, satisfy the associated access policies.

We [201] proposed a new multi-authority CP-ABE system which aims to these problems simultaneously. The system has multiple central authorities (CAs) and attribute authorities (AAs). The CAs issue identity-related keys to users but do not involve in any attribute-related operations. AAs issue attribute-related keys to users. Each AA manages a different attribute domain and operates independently from other AAs. A party may join the system to be an AA by simply registering itself to the CAs, and then publishing its attribute-related public parameters. In the proposed system, no authority can independently decrypt any ciphertext. The system is adaptively secure in the standard model which captures adaptive authority corruption. Its access policy can be any monotone access structure and the system supports large attribute universe. The efficiency of the system is also comparable to the corresponding single-authority CP-ABE system [181], which is regarded as the “state-of-the-art” single-authority CP-ABE system.

Figure 3.9 shows the architecture of the multi-authority CPABE system. The system has  $D$  Central Authorities,  $CA_1, \dots, CA_D$ , and  $K$  Attribute Authorities,

$$AA_1, \dots, AA_K.$$

Each AA manages a different domain of attributes (e.g.,  $AA_1$  manages  $U_1$ , and so on). When a user joins the system, each CA issues an identity-related key to the user. Then the user obtains an attribute-related key corresponding to the attributes that the user entitled from an AA (e.g., UNIV.X Registry). In practice, one may imagine that there could be multiple CAs run by different organizations while all of them are governed under some ordinance made by the government, then universities and companies can join the system as AAs. Each AA manages its own attribute domain and the AAs

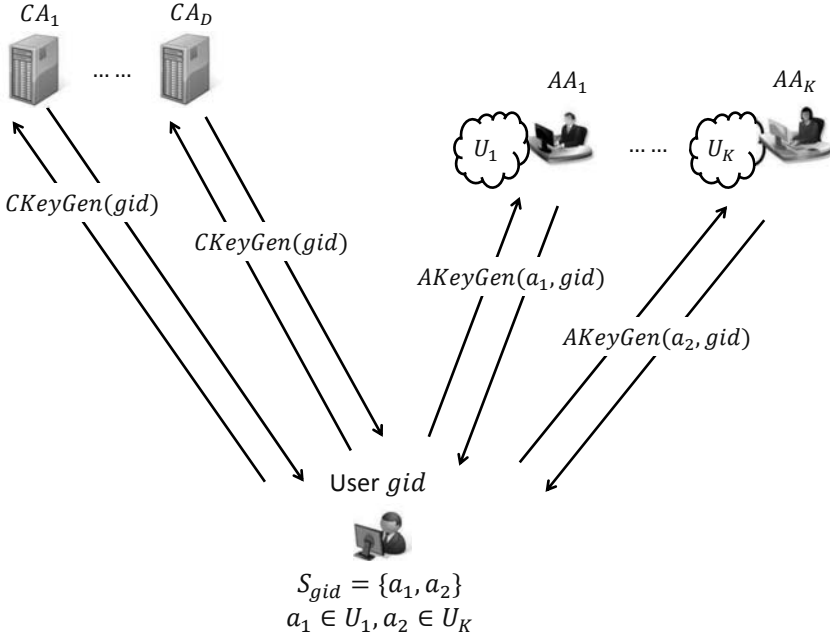


Figure 3.9: Architecture of Multi-Authority CP-ABE

operate independently from each other. The trust on the CAs by the users in the system can also be alleviated as it is unlikely to have all the CAs collude if some appropriate governmental policies and business measures are put into place to govern the practice of the CAs.

### 3.4.1 Security Models

Similar to [181, 184], the system is constructed over composite order groups. Let  $\mathcal{G}$  be the group generator, which takes a security parameter  $\lambda$  and outputs

$$(p_1, p_2, p_3, G, G_T, e)$$

where  $p_1, p_2$ , and  $p_3$  are distinct primes,  $G$  and  $G_T$  are cyclic groups of order  $N = p_1 p_2 p_3$ , and  $e : G \times G \rightarrow G_T$  is a map such that: (1) (Bilinear)  $\forall g, h \in G, a, b \in$

$\mathbb{Z}_N, e(g^a, h^b) = e(g, h)^{ab}$ , (2) (NonDegenerate)  $\exists g \in G$  such that  $e(g, g)$  has order  $N$  in  $G_T$ . Assume that group operations in  $G$  and  $G_T$  as well as the bilinear map  $e$  are computable in polynomial time with respect to  $\lambda$ . Let  $G_{p_1}$ ,  $G_{p_2}$ , and  $G_{p_3}$  be the subgroups of order  $p_1$ ,  $p_2$ , and  $p_3$  in  $G$ , respectively. Note that for any  $h_i \in G_{p_i}$  and  $h_j \in G_{p_j}$  where  $i \neq j$ ,  $e(h_i, h_j) = 1$ .

The security of the system is based on the following three assumptions, which are also used by [183](for IBE) and [181] (for CP-ABE) for obtaining full security.

For an element  $T \in G$ ,  $T$  can (uniquely) be written as the product of elements of  $G_{p_1}$ ,  $G_{p_2}$ , and  $G_{p_3}$ , and they are referred to as the “ $G_{p_1}$  part of  $T$ ,” “ $G_{p_2}$  part of  $T$ ,” and “ $G_{p_3}$  part of  $T$ ,” respectively. In the assumptions below, let  $G_{p_1 p_2}$  and  $G_{p_1 p_3}$  be the subgroups of order  $p_1 p_2$  and  $p_1 p_3$  in  $G$ , respectively. Similarly, an element in  $G_{p_1 p_2}$  can be written as the product of elements of  $G_{p_1}$  and  $G_{p_2}$ , and an element in  $G_{p_1 p_3}$  can be written as the product of elements of  $G_{p_1}$  and  $G_{p_3}$ .

**Assumption 1 (Subgroup decision problem for 3 primes).** [183] Given a group generator  $\mathcal{G}$ , define the following distribution:  $\mathbb{G} = (N = p_1 p_2 p_3, G, G_T, e) \xleftarrow{R} \mathcal{G}$ ,  $g \xleftarrow{R} G_{p_1}$ ,  $X_3 \xleftarrow{R} G_{p_3}$ ,  $D = (\mathbb{G}, g, X_3)$ ,  $T_1 \xleftarrow{R} G_{p_1 p_2}$ ,  $T_2 \xleftarrow{R} G_{p_1}$ . The advantage of an algorithm  $\mathcal{A}$  in breaking Assumption 1 is:

$$Adv1_{\mathcal{G}, \mathcal{A}}(\lambda) := |\Pr[\mathcal{A}(D, T_1) = 1] - \Pr[\mathcal{A}(D, T_2) = 1]|.$$

**Definition 3.4.1.**  $\mathcal{G}$  satisfies Assumption 1 if  $Adv1_{\mathcal{G}, \mathcal{A}}(\lambda)$  is a negligible function of  $\lambda$  for any polynomial time algorithm  $\mathcal{A}$ .

**Assumption 2.** [183] Given  $\mathcal{G}$ , define the following distribution:  $\mathbb{G} = (N = p_1 p_2 p_3, G, G_T, e) \xleftarrow{R} \mathcal{G}$ ,  $g, X_1 \xleftarrow{R} G_{p_1}$ ,  $X_2, Y_2 \xleftarrow{R} G_{p_2}$ ,  $X_3, Y_3 \xleftarrow{R} G_{p_3}$ ,  $D = (\mathbb{G}, g, X_1 X_2, X_3, Y_2 Y_3)$ ,  $T_1 \xleftarrow{R} G$ ,  $T_2 \xleftarrow{R} G_{p_1 p_3}$ . The advantage of an algorithm  $\mathcal{A}$  against Assumption 2 is:

$$Adv2_{\mathcal{G}, \mathcal{A}}(\lambda) := |\Pr[\mathcal{A}(D, T_1) = 1] - \Pr[\mathcal{A}(D, T_2) = 1]|.$$

**Definition 3.4.2.**  $\mathcal{G}$  satisfies Assumption 2 if  $\text{Adv}2_{\mathcal{G},\mathcal{A}}(\lambda)$  is a negligible function of  $\lambda$  for any polynomial time algorithm  $\mathcal{A}$ .

**Assumption 3.** [183] Given  $\mathcal{G}$ , define the following distribution:  $\mathbb{G} = (N = p_1 p_2 p_3, G, G_T, e) \xleftarrow{R} \mathcal{G}$ ,  $\alpha, s \xleftarrow{R} \mathbb{Z}_N$ ,  $g \xleftarrow{R} G_{p_1}$ ,  $X_2, Y_2, Z_2 \xleftarrow{R} G_{p_2}$ ,  $X_3 \xleftarrow{R} G_{p_3}$ ,  $D = (\mathbb{G}, g, g^\alpha X_2, X_3, g^s Y_2, Z_2)$ ,  $T_1 = e(g, g)^{\alpha s}$ ,  $T_2 \xleftarrow{R} G_T$ . The advantage of an algorithm  $\mathcal{A}$  against Assumption 3 is

$$\text{Adv}3_{\mathcal{G},\mathcal{A}}(\lambda) := |\Pr[\mathcal{A}(D, T_1) = 1] - \Pr[\mathcal{A}(D, T_2) = 1]|.$$

**Definition 3.4.3.**  $\mathcal{G}$  satisfies Assumption 3 if  $\text{Adv}3_{\mathcal{G},\mathcal{A}}(\lambda)$  is a negligible function of  $\lambda$  for any polynomial time algorithm  $\mathcal{A}$ .

**Notations.** There are three sets of entities in a multi-authority ciphertext-policy attribute-based encryption (MA-CP-ABE) system: (1) central authorities ( $CAs$ ), (2) attribute authorities ( $AAs$ ) and, (3) users. Let  $CA_1, \dots, CA_D$  be central authorities and  $\mathbb{D} = \{1, \dots, D\}$  the index set of the  $CAs$ , that is, using  $d \in \mathbb{D}$  to denote the index of central authority  $CA_d$ . Let  $AA_1, \dots, AA_K$  be attribute authorities and  $\mathbb{K} = \{1, \dots, K\}$  the index set of the  $AAs$ . Each user has a global identifier denoted as  $gid$ . The  $CAs$  are responsible for issuing keys to users according to their global identifiers. The  $AAs$  are responsible for issuing keys corresponding to attributes, and each  $AA$  manages a different attribute domain (e.g.,  $AA_i$  manages attributes for a university registry and  $AA_j$  manages attributes for a company registry, etc.). Let  $U_k$  be the attribute domain managed by  $AA_k$  where  $U_i \cap U_j = \emptyset$  for all  $i \neq j \in \mathbb{K}$ , and  $U = \bigcup_{k=1}^K U_k$  be the attribute universe.

**Definition.** A MA-CPABE system consists of the following seven algorithms:

**GlobalSetup**( $\lambda$ )  $\rightarrow$  (**GPk**): The algorithm takes as input the security parameter  $\lambda$  and outputs the global public parameter **GPk** of the system.

$\text{CASetup}(\text{GPK}, d) \rightarrow (\text{CPK}_d, \text{CAPK}_d, \text{CMSK}_d)$ : Each  $CA_d$  runs the algorithm with  $\text{GPK}$  and its index  $d$  as input, and produces master secret key  $\text{CMSK}_d$  and public parameters  $(\text{CPK}_d, \text{CAPK}_d)$ .  $\text{CAPK}_d$  will be used by  $AA$ s only.

$\text{AASetup}(\text{GPK}, k, U_k) \rightarrow (\text{APK}_k, \text{ACPK}_k, \text{AMSK}_k)$ : Each  $AA_k$  runs the algorithm with  $\text{GPK}$ , its index  $k$  and its attribute domain  $U_k$  as input, and produces master secret key  $\text{AMSK}_k$  and public parameters  $(\text{APK}_k, \text{ACPK}_k)$ .  $\text{ACPK}_k$  will be used by  $CA$ s only.

$\text{Encrypt}(M, \mathbb{A}, \text{GPK}, \{\text{CPK}_d | d \in \mathbb{D}\}, \{\text{APK}_k\}) \rightarrow CT$ : The algorithm takes as input a message  $M$ , an access policy  $\mathbb{A}$  defined over the attribute universe  $U$ , the global public parameter  $\text{GPK}$ ,  $CA$ s' public parameters  $\{\text{CPK}_d | d \in \mathbb{D}\}$ , and the related  $AA$ s' public parameters  $\{\text{APK}_k\}$ . It outputs a ciphertext  $CT$  which contains the access policy  $\mathbb{A}$ .

$\text{CKeyGen}(gid, \text{GPK}, \{\text{ACPK}_k | k \in \mathbb{K}\}, \text{CMSK}_d) \rightarrow (\text{ucsk}_{gid,d}, \text{ucpk}_{gid,d})$ : When a user with global identifier  $gid$  visits  $CA_d$  for obtaining a key,  $CA_d$  runs the algorithm, which takes as input  $gid$ ,  $\text{GPK}$ ,  $\{\text{ACPK}_k | k \in \mathbb{K}\}$ , and  $CA_d$ 's master secret key  $\text{CMSK}_d$ . It outputs a **user-central-key**  $(\text{ucsk}_{gid,d}, \text{ucpk}_{gid,d})$ , where  $\text{ucpk}_{gid,d}$  is called **user-central-public-key**.

$\text{AKeyGen}(att, \{\text{ucpk}_{gid,d} | d \in \mathbb{D}\}, \text{GPK}, \{\text{CAPK}_d | d \in \mathbb{D}\}, \text{AMSK}_k) \rightarrow \text{uask}_{att,gid} \text{ or } \perp$ : When a user requests a secret key for attribute  $att$  from  $AA_k$ ,  $AA_k$  runs the algorithm, which takes as input  $att$ ,  $\{\text{ucpk}_{gid,d} | d \in \mathbb{D}\}$ ,  $\text{GPK}$ ,  $\{\text{CAPK}_d | d \in \mathbb{D}\}$  and  $\text{AMSK}_k$ . If all  $\text{ucpk}_{gid,d}$ s are valid, the algorithm outputs a **user-attribute-key**  $\text{uask}_{att,gid}$ , otherwise it outputs  $\perp$ . For a user  $gid$  with attribute set  $S_{gid}$ , the user's **decryption-key** is defined as

$$\text{DK}_{gid} = (\{\text{ucsk}_{gid,d}, \text{ucpk}_{gid,d} | d \in \mathbb{D}\}, \{\text{uask}_{att,gid} | att \in S_{gid}\}).$$

$\text{Decrypt}(CT, \text{GPK}, \{\text{APK}_k\}, \text{DK}_{gid}) \rightarrow M$  or  $\perp$ : The algorithm takes as input a ciphertext  $CT$  associated with access policy  $\mathbb{A}$ ,  $\text{GPK}$ , the related attribute authorities' public parameters  $\{\text{APK}_k\}$ , and a decryption-key  $\text{DK}_{gid}$  with attribute set  $S_{gid}$ . If  $S_{gid}$  satisfies the access policy  $\mathbb{A}$ , the algorithm outputs the message  $M$ , otherwise it outputs  $\perp$ .

**Security Model** : The security of MA-CP-ABE is defined by the following game run between a challenger  $\mathcal{B}$  and an adversary  $\mathcal{A}$ .  $\mathcal{A}$  can corrupt  $CA$ s and  $AA$ s by specifying  $\mathbb{K}_c \subset \mathbb{K}$  and  $\mathbb{D}_c \subset \mathbb{D}$  after seeing the public parameters<sup>6</sup>, where  $\mathbb{D} \setminus \mathbb{D}_c \neq \emptyset$  and  $\mathbb{K} \setminus \mathbb{K}_c \neq \emptyset$ . Without loss of generality, we assume that  $\mathcal{A}$  corrupts all  $CA$ s but one, i.e.,  $|\mathbb{D} \setminus \mathbb{D}_c| = 1$ .

**Setup.**

- $\text{GlobalSetup}, \text{CASetup}(\text{GPK}, d)$  ( $d = 1, \dots, D$ ) and  $\text{AASetup}(\text{GPK}, k, U_k)$  ( $k = 1, \dots, K$ ) are run by the challenger  $\mathcal{B}$ .  $\text{GPK}, \{\text{CPK}_d, \text{CAPK}_d | d \in \mathbb{D}\}$  and  $\{\text{APK}_k, \text{ACPK}_k | k \in \mathbb{K}\}$  are given to the adversary  $\mathcal{A}$ .
- $\mathcal{A}$  specifies an index  $d^* \in \mathbb{D}$  as the only uncorrupted  $CA$  and specifies a set  $\mathbb{K}_c \subset \mathbb{K}$  of  $AA$ s to be corrupted where  $\mathbb{K} \setminus \mathbb{K}_c \neq \emptyset$ . Let  $\mathbb{D}_c = \mathbb{D} \setminus \{d^*\}$ .  $\{\text{CMSK}_d | d \in \mathbb{D}_c\}$  and  $\{\text{AMSK}_k | k \in \mathbb{K}_c\}$  are given to  $\mathcal{A}$ .

**Key Query Phase 1.** User-central-key and user-attribute-key can be obtained by querying the following oracles:

$\text{CKQ}(gid, d)$  where  $d = d^*$ :  $\mathcal{A}$  queries with a pair  $(gid, d)$ , where  $gid$  is a global identifier and  $d = d^*$ , and obtains the corresponding user-central-key  $(\text{ucsk}_{gid, d^*}, \text{ucpk}_{gid, d^*})$ .

$\text{AKQ}(att, \{\text{ucpk}_{gid, d} | d \in \mathbb{D}\}, k)$  where  $k \in \mathbb{K} \setminus \mathbb{K}_c$ :  $\mathcal{A}$  queries with  $(att, \{\text{ucpk}_{gid, d} | d \in \mathbb{D}\}, k)$ , where  $k \in \mathbb{K} \setminus \mathbb{K}_c$  is the index of an uncorrupted

<sup>6</sup>This is stronger than the static corruption model used in [78, 79, 184], where the adversary has to specify the authorities to corrupt before seeing the public parameters. But on the other aspect, it is weaker than the model in [184], where the corrupted authorities are set by the adversary.

$AA$ ,  $\{\text{ucpk}_{gid,d} | d \in \mathbb{D}\}$  are  $gid$ 's user-central-public-keys, and  $att$  is an attribute in  $U_k$ . The oracle returns a user-attribute-key  $\text{uask}_{att,gid}$  or  $\perp$  if  $\{\text{ucpk}_{gid,d}\}$  is invalid.

**Challenge Phase.**  $\mathcal{A}$  submits two equal-length messages  $M_0$  and  $M_1$ , and an access policy  $\mathbb{A}$ .  $\mathcal{B}$  flips a random coin  $\beta \in \{0, 1\}$  and sends to  $\mathcal{A}$  an encryption of  $M_\beta$  under  $\mathbb{A}$ .

**Phase 2.**  $\mathcal{A}$  further queries as in **Key Query Phase 1**.

**Guess.**  $\mathcal{A}$  submits a guess  $\beta'$  for  $\beta$ .

For a  $gid$ , the related attribute set is defined as

$$S_{gid} = \{att \mid \text{AKQ}(att, \{\text{ucpk}_{gid,d} | d \in \mathbb{D}\}, k) \text{ is made by } \mathcal{A}\}.$$

$\mathcal{A}$  wins the game if  $\beta' = \beta$  under the **restriction** that there is no  $S_{gid}$  such that  $S_{gid} \cup (\bigcup_{k_c \in \mathbb{K}_c} U_{k_c})$  can satisfy the challenge access policy  $\mathbb{A}$ . The advantage of  $\mathcal{A}$  is defined as  $|\Pr[\beta = \beta'] - 1/2|$ .

**Definition 3.4.4.** An MA-CP-ABE system is secure if for all polynomial-time adversary  $\mathcal{A}$  in the game above, the advantage of  $\mathcal{A}$  is negligible.

**Remarks :** It is assumed that a user with global identifier  $gid$  requests for the central key from each  $CA_d$  only once, i.e., for each  $gid$  there is only one set of user-central-keys,  $\{\text{ucpk}_{gid,d} | d \in \mathbb{D}\}$ . This is not a restriction, but can help simplify the system description. Using obscure notations such as  $\text{ucpk}_{gid,d,t}$  and  $S_{gid,d,t}$  where  $t$  is a time stamp can remove this assumption. In the security model above,  $\mathcal{A}$  has the master secret keys  $\{\text{CMSK}_d | d \in \mathbb{D}_c\}$ , so the user only needs to query  $\text{CKQ}(gid, d^*)$  for getting  $(\text{ucsk}_{gid,d^*}, \text{ucpk}_{gid,d^*})$ , and he/she can generate  $\{(\text{ucsk}_{gid,d}, \text{ucpk}_{gid,d}) | d \in \mathbb{D}_c\}$  if they are needed for querying  $\text{AKQ}$ .



### 3.4.2 Construction

**GlobalSetup**( $\lambda$ )  $\rightarrow$  (**GPK**): Let  $G$  be a bilinear group of order  $N = p_1 p_2 p_3$  (3 distinct primes), and  $G_{p_i}$  be the subgroup of order  $p_i$  in  $G$ . The algorithm randomly chooses  $g, h \in G_{p_1}$ . Let  $X_3$  be a generator of  $G_{p_3}$ .

The global public parameter is published as **GPK** =  $(N, g, h, X_3, \Sigma_{sign})$ , where  $\Sigma_{sign} = (\text{KeyGen}, \text{Sign}, \text{Verify})$  is the description of a UF-CMA secure signature scheme.

**CASetup**(**GPK**,  $d$ )  $\rightarrow$  (**CPK** $_d$ , **CAPK** $_d$ , **CMSK** $_d$ ):  $CA_d$  runs the **KeyGen** algorithm of  $\Sigma_{sign}$  to generate sign key pair (**SignKey** $_d$ , **VerifyKey** $_d$ ), and chooses a random exponent  $\alpha_d \in \mathbb{Z}_N$ .

$CA_d$  publishes its public parameter **CPK** $_d = e(g, g)^{\alpha_d}$ , **CAPK** $_d = \text{VerifyKey}_d$ .

$CA_d$  sets its master secret key **CMSK** $_d = (\alpha_d, \text{SignKey}_d)$ .

**AASetup**(**GPK**,  $k, U_k$ )  $\rightarrow$  (**APK** $_k$ , **ACPK** $_k$ , **AMSK** $_k$ ): For each  $att \in U_k$ ,  $AA_k$  randomly chooses  $s_{att} \in \mathbb{Z}_N$  and sets  $T_{att} = g^{s_{att}}$ . For each  $d \in \mathbb{D}$ ,  $AA_k$  randomly chooses  $v_{k,d} \in \mathbb{Z}_N$  and sets  $V_{k,d} = g^{v_{k,d}}$ .

$AA_k$  publishes its public parameter **APK** $_k = \{T_{att} | att \in U_k\}$ , **ACPK** $_k = \{V_{k,d} | d \in \mathbb{D}\}$ .

$AA_k$  sets its master secret key **AMSK** $_k = (\{s_{att} | att \in U_k\}, \{v_{k,d} | d \in \mathbb{D}\})$ .

**Encrypt**( $M, \mathbb{A} = (A, \rho)$ , **GPK**,  $\{\text{CPK}_d | d \in \mathbb{D}\}$ ,  $\{\text{APK}_k\}$ )  $\rightarrow CT$ :  $M$  is the message to be encrypted,  $\mathbb{A}$  is the access policy which is expressed by an LSSS matrix  $(A, \rho)$ , where  $A$  is an  $l \times n$  matrix and  $\rho$  maps each row  $A_x$  of  $A$  to an attribute  $\rho(x)$ . Here it is required that  $\rho$  will not map two different rows to a same attribute. The algorithm chooses a random vector  $\vec{v} = (s, v_2, \dots, v_n) \in \mathbb{Z}_N^n$ , and for each  $x \in \{1, 2, \dots, l\}$ , it randomly picks  $r_x \in \mathbb{Z}_N$ . Let  $A_x \cdot \vec{v}$  be the inner product of

the  $x^{th}$  row of  $A$  and the vector  $\vec{v}$ . The ciphertext is

$$C = M \cdot \prod_{d=1}^D e(g, g)^{\alpha_d \cdot s}, \quad C' = g^s,$$

$$\{C_x = h^{A_x \cdot \vec{v}} T_{\rho(x)}^{-r_x}, C'_x = g^{r_x} \mid x \in \{1, 2, \dots, l\}\}$$

along with the access policy  $\mathbb{A} = (A, \rho)$ .

**CKeyGen**( $gid$ ,  $\text{GPK}$ ,  $\{V_{k,d} \mid k \in \mathbb{K}\}$ ,  $\text{CMSK}_d$ )  $\rightarrow$  ( $\text{ucsk}_{gid,d}$ ,  $\text{ucpk}_{gid,d}$ ): When a user submits his  $gid$  to  $CA_d$  to request the user-central-key,  $CA_d$  randomly chooses  $r_{gid,d} \in \mathbb{Z}_N$  and  $R_{gid,d}, R'_{gid,d} \in G_{p_3}$ , then sets

$$\text{ucsk}_{gid,d} = g^{\alpha_d} h^{r_{gid,d}} R_{gid,d}, \quad L_{gid,d} = g^{r_{gid,d}} R'_{gid,d}.$$

For  $k = 1$  to  $K$ ,  $CA_d$  randomly picks  $R_{gid,d,k} \in G_{p_3}$  and computes

$$\Gamma_{gid,d,k} = V_{k,d}^{r_{gid,d}} R_{gid,d,k}.$$

$CA_d$  computes  $\sigma_{gid,d} = \text{Sign}(\text{SignKey}_d, gid \parallel d \parallel L_{gid,d} \parallel \Gamma_{gid,d,1} \parallel \dots \parallel \Gamma_{gid,d,K})$ .

Let  $\text{ucpk}_{gid,d} = (gid, d, L_{gid,d}, \{\Gamma_{gid,d,k} \mid k \in \mathbb{K}\}, \sigma_{gid,d})$ .

**AKeyGen**( $att$ ,  $\{\text{ucpk}_{gid,d} \mid d \in \mathbb{D}\}$ ,  $\text{GPK}$ ,  $\{\text{VerifyKey}_d \mid d \in \mathbb{D}\}$ ,  $\text{AMSK}_k$ )  $\rightarrow$   $\text{uask}_{att,gid}$  or  $\perp$ : When a user submits his  $\{\text{ucpk}_{gid,d} \mid d \in \mathbb{D}\}$  to  $AA_k$  to request the user-attribute-key for attribute  $att \in U_k$ ,

1. For  $d = 1$  to  $D$ ,  $AA_k$  parses  $\text{ucpk}_{gid,d}$  into  $(gid, d, L_{gid,d}, \{\Gamma_{gid,d,k} \mid k \in \mathbb{K}\}, \sigma_{gid,d})$  and checks whether

$$valid \leftarrow \text{Verify}(\text{VerifyKey}_d, gid \parallel d \parallel L_{gid,d} \parallel \Gamma_{gid,d,1} \parallel \dots \parallel \Gamma_{gid,d,K}, \sigma_{gid,d}) \quad (3.1)$$

$$e(g, \Gamma_{gid,d,k}) = e(V_{k,d}, L_{gid,d}) \neq 1. \quad (3.2)$$

If fails,  $AA_k$  outputs  $\perp$  to the user to imply that the submitted  $\{\text{ucpk}_{gid,d} | d \in \mathbb{D}\}$  are invalid.

2. For  $d = 1$  to  $D$ ,  $AA_k$  randomly picks  $R'_{att,gid,d} \in G_{p_3}$ , and sets

$$\text{uask}_{att,gid,d} = (\Gamma_{gid,d,k})^{s_{att}/v_{k,d}} R'_{att,gid,d}. \quad (3.3)$$

Note that

$$\begin{aligned} \text{uask}_{att,gid,d} &= (\Gamma_{gid,d,k})^{s_{att}/v_{k,d}} R'_{att,gid,d} \\ &= (V_{k,d}^{r_{gid,d}} R_{gid,d,k})^{s_{att}/v_{k,d}} R'_{att,gid,d} \\ &= (g^{v_{k,d} \cdot r_{gid,d}} R_{gid,d,k})^{s_{att}/v_{k,d}} R'_{att,gid,d} \\ &= T_{att}^{r_{gid,d}} (R_{gid,d,k})^{s_{att}/v_{k,d}} R'_{att,gid,d} \end{aligned}$$

As  $(R_{gid,d,k})^{s_{att}/v_{k,d}} R'_{att,gid,d}$  is in  $G_{p_3}$  and  $R'_{att,gid,d}$  is randomly chosen, we can write

$$\text{uask}_{att,gid,d} = T_{att}^{r_{gid,d}} R_{att,gid,d}. \quad (3.4)$$

Without knowing the value of  $r_{gid,d}$ , by running (3.3),  $AA_k$  can compute the value as (3.4).

3.  $AA_k$  outputs user-attribute-key  $\text{uask}_{att,gid}$  to user where

$$\begin{aligned} \text{uask}_{att,gid} &= \prod_{d=1}^D \text{uask}_{att,gid,d} = \prod_{d=1}^D T_{att}^{r_{gid,d}} R_{att,gid,d} \\ &= T_{att}^{\sum_{d=1}^D r_{gid,d}} \prod_{d=1}^D R_{att,gid,d} \\ &= T_{att}^{\sum_{d=1}^D r_{gid,d}} R_{att,gid} \end{aligned} \quad (3.5)$$

$\text{Decrypt}(CT, \text{GPK}, \{\text{APK}_k\}, \text{DK}_{gid}) \rightarrow M$ : The ciphertext  $CT$  is parsed into  $\langle C, C', \{C_x, C'_x | x \in \{1, 2, \dots, l\}\}, \mathbb{A} = (A, \rho) \rangle$ , and the decryption-key  $\text{DK}_{gid}$  is parsed

into  $(\{\text{ucsk}_{gid,d}, \text{ucpk}_{gid,d} | d \in \mathbb{D}\}, \{\text{uask}_{att,gid} | att \in S_{gid}\})$ .

The algorithm computes

$$\begin{aligned}
 \bullet \text{ucsk}_{gid} &= \prod_{d=1}^D \text{ucsk}_{gid,d} = g^{\sum_{d=1}^D \alpha_d} h^{\sum_{d=1}^D r_{gid,d}} \prod_{d=1}^D R_{gid,d} = \\
 &g^\alpha h^{r_{gid}} R_{gid}, \\
 \text{with } \alpha &= \sum_{d=1}^D \alpha_d, r_{gid} = \sum_{d=1}^D r_{gid,d} \text{ and } R_{gid} = \prod_{d=1}^D R_{gid,d}. \\
 \bullet L_{gid} &= \prod_{d=1}^D L_{gid,d} = g^{\sum_{d=1}^D r_{gid,d}} \prod_{d=1}^D R'_{gid,d} = g^{r_{gid}} R'_{gid}, \\
 \text{with } R'_{gid} &= \prod_{d=1}^D R'_{gid,d}.
 \end{aligned}$$

Note that  $\forall att \in S_{gid}, \text{uask}_{att,gid} = T_{att}^{\sum_{d=1}^D r_{gid,d}} R_{att,gid} = T_{att}^{r_{gid}} R_{att,gid}$ .

If  $S_{gid}$  satisfies the access policy  $(A, \rho)$ , the algorithm computes constants  $\omega_x \in \mathbb{Z}_N$  such that  $\sum_{\rho(x) \in S_{gid}} \omega_x A_x = (1, 0, \dots, 0)$ . Then it computes

$$e(C', \text{ucsk}_{gid}) / \prod_{\rho(x) \in S_{gid}} (e(C_x, L_{gid}) \cdot e(C'_x, \text{uask}_{\rho(x),gid}))^{\omega_x} = e(g, g)^{\alpha s}.$$

As  $C = M \cdot \prod_{d=1}^D e(g, g)^{\alpha_d \cdot s} = M \cdot e(g, g)^{s \sum_{d=1}^D \alpha_d} = M \cdot e(g, g)^{s \alpha}$ ,  $M$  can be recovered from  $C / e(g, g)^{\alpha s}$ .

In the above system, an attribute is required to appear at most once in an LSSS matrix  $(A, \rho)$ . This restriction is crucial to the security proof. Such a system is called a one-use system, and it can be extended to a multi-use system by using the encoding technique [181].

### 3.4.3 Security Analysis

Let  $\Pi$  denote the main construction, we modify  $\Pi$  to  $\Pi'$  as follows:

- In the AKeyGen algorithm, it outputs  $\text{uask}_{att,gid} = \{\text{uask}_{att,gid,d} | d \in \mathbb{D}\}$

rather than  $\text{uask}_{att,gid} = \prod_{d=1}^D \text{uask}_{att,gid,d}$ . i.e.,  $gid$ 's decryption-key is

$$\begin{aligned}
 \text{DK}_{gid} &= (\{\text{ucsk}_{gid,d}, \text{ucpk}_{gid,d} | d \in \mathbb{D}\}, \{\text{uask}_{att,gid} | att \in S_{gid}\}) \\
 &= (\{\text{ucsk}_{gid,d}, \text{ucpk}_{gid,d} | d \in \mathbb{D}\}, \{\{\text{uask}_{att,gid,d} | d \in \mathbb{D}\} | att \in S_{gid}\}) \\
 &= (\{\text{ucsk}_{gid,d}, \text{ucpk}_{gid,d} | d \in \mathbb{D}\}, \{\{\text{uask}_{att,gid,d} | att \in S_{gid}\} | d \in \mathbb{D}\}) \\
 &= \{(\text{ucsk}_{gid,d}, \text{ucpk}_{gid,d}, \{\text{uask}_{att,gid,d} | att \in S_{gid}\}) | d \in \mathbb{D}\} \\
 &= \{\text{usk}_{gid,d} | d \in \mathbb{D}\}
 \end{aligned}$$

where  $\text{usk}_{gid,d} = (\text{ucsk}_{gid,d}, \text{ucpk}_{gid,d}, \{\text{uask}_{att,gid,d} | att \in S_{gid}\})$  is called  $gid$ 's **user-key** related to  $d$ .

- In the Decrypt algorithm,

1. For  $d = 1$  to  $D$ , the algorithm uses  $\text{usk}_{gid,d}$  to reconstruct  $e(g, g)^{\alpha_{d^*} s}$ :

$$e(C', \text{ucsk}_{gid,d}) / \prod_{\rho(x) \in S_{gid}} (e(C_x, L_{gid,d}) \cdot e(C'_x, \text{uask}_{\rho(x),gid,d}))^{\omega_x} = e(g, g)^{\alpha_{d^*} s}. \quad (3.6)$$

2. The algorithm recovers  $M$  by

$$M = C / \prod_{d=1}^D e(g, g)^{\alpha_{d^*} s}. \quad (3.7)$$

Note that the user and the attacker will get more information in  $\Pi'$ , the security of  $\Pi'$  will imply the security of  $\Pi$ . We show the security of  $\Pi'$  in the following: In the security model,  $CA_{d^*}$  is the only uncorrupted central authority and no  $S_{gid} \cup (\bigcup_{k_c \in \mathbb{K}_c} U_{k_c})$  can satisfy the challenge access policy. It means that the adversary could not request keys to form a  $\text{usk}_{gid,d^*}$  to reconstruct  $e(g, g)^{\alpha_{d^*} s}$ . In our proof, the challenger will respond the adversary as in real attack for all key queries related to  $d \neq d^*$ . On the key queries related to  $d^*$ , we use the proof technique [181] to provide the answers. Before we give our proof, we need to define two additional structures: semi-functional ciphertexts and keys. We choose random values  $z_{att} \in \mathbb{Z}_N$  associated to the attributes.

**Semi-functional Ciphertext:** A semi-functional ciphertext is formed as follows. Let  $g_2$  denote a generator of  $G_{p_2}$  and  $c$  a random exponent modulo  $N$ . Besides the random vector  $\vec{v} = (s, v_2, \dots, v_n)$  and the random values  $\{r_x | x \in \{1, 2, \dots, l\}\}$ , we also choose a random vector  $\vec{u} = (u_1, u_2, \dots, u_n) \in \mathbb{Z}_N^n$  and random values  $\{\gamma_x \in \mathbb{Z}_N | x \in \{1, 2, \dots, l\}\}$ . Then:

$$C' = g^s g_2^c, \{C_x = h^{A_x \cdot \vec{v}} T_{\rho(x)}^{-r_x} g_2^{A_x \vec{u} + \gamma_x z_{\rho(x)}}, C'_x = g^{r_x} g_2^{-\gamma_x} \mid x \in \{1, 2, \dots, l\}\}.$$

**Semi-functional Key:** For a  $gid$ , a semi-functional user-key  $\text{usk}_{gid,d^*}$  has two possible forms. Exponents  $r_{gid,d^*}, \delta, b \in \mathbb{Z}_N$ ,  $\{w_{k,d^*} \in \mathbb{Z}_N | k \in \mathbb{K}\}$ , and elements  $R_{gid,d^*}, R'_{gid,d^*} \in G_{p_3}$ ,  $\{R_{att,gid,d^*} \in G_{p_3} | att \in S_{gid}\}$ ,  $\{R_{gid,d^*,k} \in G_{p_3} | k \in \mathbb{K}\}$  are chosen randomly.

- Type 1:

The user-central-key ( $\text{ucsk}_{gid,d^*}, \text{ucpk}_{gid,d^*}$ ) is formed as

$$\text{ucsk}_{gid,d^*} = g^{\alpha_{d^*}} h^{r_{gid,d^*}} R_{gid,d^*} g_2^\delta, \quad L_{gid,d^*} = g^{r_{gid,d^*}} R'_{gid,d^*} g_2^b,$$

$$\Gamma_{gid,d^*,k} = V_{k,d^*}^{r_{gid,d^*}} R_{gid,d^*,k} g_2^{bw_{k,d^*}} \quad (k = 1, 2, \dots, K),$$

$$\sigma_{gid,d^*} = \text{Sign}(\text{SignKey}_{d^*}, gid || d^* || L_{gid,d^*} || \Gamma_{gid,d^*,1} || \dots || \Gamma_{gid,d^*,K}),$$

$$\text{ucpk}_{gid,d^*} = (gid, d^*, L_{gid,d^*}, \{\Gamma_{gid,d^*,k} | k \in \mathbb{K}\}, \sigma_{gid,d^*}).$$

$\forall att \in S_{gid}$ , the derived  $\text{uask}_{att,gid,d^*}$  is formed as

$$\text{uask}_{att,gid,d^*} = T_{att}^{r_{gid,d^*}} R_{att,gid,d^*} g_2^{bz_{att}}.$$

- Type 2:

The user-central-key ( $\text{ucsk}_{gid,d^*}, \text{ucpk}_{gid,d^*}$ ) is formed as

$$\begin{aligned} \text{ucsk}_{gid,d^*} &= g^{\alpha_{d^*}} h^{r_{gid,d^*}} R_{gid,d^*} g_2^\delta, \quad L_{gid,d^*} = g^{r_{gid,d^*}} R'_{gid,d^*}, \\ \Gamma_{gid,d^*,k} &= V_{k,d^*}^{r_{gid,d^*}} R_{gid,d^*,k} (k = 1, 2, \dots, K), \\ \sigma_{gid,d^*} &= \text{Sign}(\text{SignKey}_{d^*}, gid || d^* || L_{gid,d^*} || \Gamma_{gid,d^*,1} || \dots || \Gamma_{gid,d^*,K}), \\ \text{ucpk}_{gid,d^*} &= (gid, d^*, L_{gid,d^*}, \{\Gamma_{gid,d^*,k} | k \in \mathbb{K}\}, \sigma_{gid,d^*}). \end{aligned}$$

$\forall att \in S_{gid}$ , the derived  $\text{uask}_{att,gid,d^*}$  is formed as

$$\text{uask}_{att,gid,d^*} = T_{att}^{r_{gid,d^*}} R_{att,gid,d^*}.$$

Note that both the semi-functional user-keys of types 1 and 2 satisfy (3.1) and (3.2), and that type 2 is a special case of type 1 with  $b = 0$ .

When a normal  $\text{usk}_{gid,d^*}$  and a semi-functional ciphertext, or a semi-functional  $\text{usk}_{gid,d^*}$  and a normal ciphertext, are used in computation (3.6),  $e(g, g)^{\alpha_{d^*} s}$  is available, and this value could be used in the computation (3.7). When a semi-functional  $\text{usk}_{gid,d^*}$  and a semi-functional ciphertext are used in computation (3.6), the result is  $e(g, g)^{\alpha_{d^*} s} \cdot e(g_2, g_2)^{c\delta - bu_1}$ . The additional term  $e(g_2, g_2)^{c\delta - bu_1}$  will hinder the computation (3.7). We call a semi-functional user-key of type 1 *nominally semi-functional* if  $c\delta - bu_1 = 0$ .

The security of  $\Pi'$  relies on Assumptions 1, 2, and 3. We use a hybrid argument over a sequence of games. The first game **Game<sub>Real</sub>** is the real security game. In the final game **Game<sub>Final</sub>**, all user-keys related to  $d^*$ ,  $\{\text{usk}_{gid,d^*}\}$ , are semi-functional of type 2 and the ciphertext is a semi-functional encryption of a random message, independent of the two messages provided by  $\mathcal{A}$ .

**Game<sub>Real</sub>** : The challenge ciphertext is normal. All CKQs are answered with normal user-central-key. All AKQs are answered with user-attribute-key generated by running the normal AKeyGen algorithm.

**Game<sub>0</sub>** : The challenge ciphertext is semi-functional. All CKQs are answered with normal user-central-key. All AKQs are answered with user-attribute-key generated by running the normal AKeyGen algorithm.

Let  $q$  denote the number of CKQ made by  $\mathcal{A}$ . For  $j$  from 1 to  $q$ , we consider the following games:

**Game<sub>j,1</sub>** : In this game, the challenge ciphertext is semi-functional. The first  $j - 1$  CKQs are answered with semi-functional user-central-key of type 2; the  $j^{th}$  CKQ is answered with semi-functional user-central-key of type 1; and the remaining CKQs are answered with normal user-central-key. All AKQs are answered with user-attribute-key generated by running the normal AKeyGen algorithm.

**Game<sub>j,2</sub>** : In this game, the challenge ciphertext is semi-functional. The first  $j - 1$  CKQs are answered with semi-functional user-central-key of type 2; the  $j^{th}$  CKQ is answered with semi-functional user-central-key of type 2; and the remaining CKQs are answered with normal user-central-key. All AKQs are answered with user-attribute-key generated by running the normal AKeyGen algorithm.

**Game<sub>Final</sub>**: In this game, the challenge ciphertext is a semi-functional encryption of a random message, independent of the two messages provided by the adversary. All CKQs are answered with semi-functional user-central-key of type 2. All AKQs are answered with user-attribute-key generated by running the normal AKeyGen algorithm, and this step has been followed in all the games.

In the proofs, we will show that the derived  $\text{uask}_{att,gid,d^*}$  is decided by the corresponding user-central-key  $(\text{ucsk}_{gid,d^*}, \text{ucpk}_{gid,d^*})$ , i.e., if  $(\text{ucsk}_{gid,d^*}, \text{ucpk}_{gid,d^*})$  is semi-functional of type 1 (respectively, type 2), then the derived  $\text{uask}_{att,gid,d^*}$  is also semi-functional of type 1 (respectively, type 2). Consequently,  $\text{usk}_{gid,d^*}$  is decided by the corresponding  $(\text{ucsk}_{gid,d^*}, \text{ucpk}_{gid,d^*})$  as well. Note that in **Game<sub>0</sub>** all user-central-keys related to  $d^*$  are normal and in **Game<sub>q,2</sub>** all user-central-keys related to



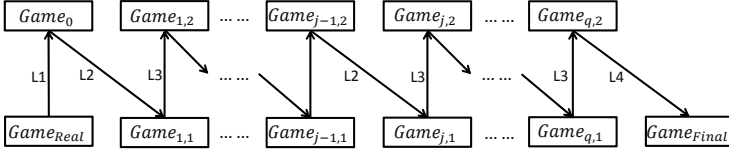


Figure 3.10: Indistinguishable games. L1 denotes Lemma 1, and so on.

$d^*$  are semi-functional of type 2. It means that in **Game**<sub>0</sub> all user-keys  $\text{usk}_{gid,d^*}$  are normal and in **Game** <sub>$q,2$</sub>  all user-keys  $\text{usk}_{gid,d^*}$  are semi-functional of type 2. We show these games are indistinguishable in the following four lemmas (see Figure 3.10).

**Lemma 3.4.5.** *Given a UF-CMA signature scheme  $\Sigma_{\text{sign}}$ , suppose a poly-time algorithm  $\mathcal{A}$  exists so that  $\mathbf{Game}_{\text{Real}}\text{Adv}_{\mathcal{A}} - \mathbf{Game}_0\text{Adv}_{\mathcal{A}} = \epsilon$ . We can construct a poly-time algorithm  $\mathcal{B}$  with advantage  $\epsilon$  against Assumption 1.*

**Lemma 3.4.6.** *Use  $\mathbf{Game}_{0,2}$  to denote  $\mathbf{Game}_0$ . Given a UF-CMA signature scheme  $\Sigma_{\text{sign}}$ , suppose a poly-time algorithm  $\mathcal{A}$  exists so that*

$$\mathbf{Game}_{j-1,2}\text{Adv}_{\mathcal{A}} - \mathbf{Game}_{j,1}\text{Adv}_{\mathcal{A}} = \epsilon.$$

*We can construct a poly-time algorithm  $\mathcal{B}$  with advantage negligibly close to  $\epsilon$  against Assumption 2.*

**Lemma 3.4.7.** *Given a UF-CMA signature scheme  $\Sigma_{\text{sign}}$ , suppose a poly-time algorithm  $\mathcal{A}$  exists so that  $\mathbf{Game}_{j,1}\text{Adv}_{\mathcal{A}} - \mathbf{Game}_{j,2}\text{Adv}_{\mathcal{A}} = \epsilon$ . We can construct a poly-time algorithm  $\mathcal{B}$  with advantage  $\epsilon$  against Assumption 2.*

**Lemma 3.4.8.** *Given a UF-CMA signature scheme  $\Sigma_{\text{sign}}$ , suppose a poly-time algorithm  $\mathcal{A}$  exists so that  $\mathbf{Game}_{q,2}\text{Adv}_{\mathcal{A}} - \mathbf{Game}_{\text{Final}}\text{Adv}_{\mathcal{A}} = \epsilon$ . We can construct a poly-time algorithm  $\mathcal{B}$  with advantage  $\frac{\epsilon}{D}$  against Assumption 3.*

**Theorem 3.4.9.** *If the signature scheme  $\Sigma_{\text{sign}}$  is UF-CMA secure and Assumptions 1, 2, and 3 hold, then our MA-CP-ABE scheme is secure.*

*Proof.* If Assumptions 1, 2 and 3 hold, and the signature scheme  $\Sigma_{sign}$  is UF-CMA secure, previous lemmas have shown that the real security game is indistinguishable from **Game**<sub>Final</sub>, in which the value of  $\beta$  is information-theoretically hidden from the adversary. Hence the adversary can not obtain a non-negligible advantage in breaking  $\Pi'$ , which implies the adversary can not obtain a non-negligible advantage in breaking our MACPABE scheme  $\Pi$ .  $\square$

### 3.4.4 Security Proofs

In the following proof, it should be noted that all AKQs are answered with user-attribute-key generated by running the **normal AKeyGen** algorithm. In the **AKeyGen** algorithm, the signature verification (3.1) used to verify  $gid$ 's  $ucpk_{gid,d^*}$ , the check (3.2), and the computation (3.3) work together to ensure that the adversary could not make use of  $ucpk_{gid,d^*}$  to construct  $ucpk_{gid,d}(d \in \mathbb{D}_c)$ .

#### Proof of Lemma 3.4.5.

*Proof.*  $\mathcal{B}$  is given  $\mathbb{G} = (N, G, G_T, e), g, X_3, T$ . It will simulate **Game**<sub>Real</sub> or **Game**<sub>0</sub> to  $\mathcal{A}$ .

$\mathcal{B}$  randomly chooses  $a \in \mathbb{Z}_N$ , then gives  $\mathcal{A}$   $\text{GPK} = \{N, g, h = g^a, X_3\}$ ,  $\mathbb{D} = \{1, 2, \dots, D\}$ ,  $\mathbb{K} = \{1, 2, \dots, K\}$ , the descriptions of  $U_1, U_2, \dots, U_K$ , and the description of secure signature scheme  $\Sigma_{sign}$ .

For  $d = 1$  to  $D$ ,  $\mathcal{B}$  generates sign key pair  $(\text{SignKey}_d, \text{VerifyKey}_d)$  and randomly chooses  $\alpha_d \in \mathbb{Z}_N$ , then gives  $\mathcal{A}$   $\text{CPK}_d = e(g, g)^{\alpha_d}$  and  $\text{CAPK}_d = \text{VerifyKey}_d$ .

For  $k = 1$  to  $K$ ,  $\mathcal{B}$  randomly chooses  $\{s_{att} \in \mathbb{Z}_N | att \in U_k\}$  and  $\{v_{k,d} \in \mathbb{Z}_N | d \in \mathbb{D}\}$ , then gives  $\mathcal{A}$   $\text{APK}_k = \{T_{att} = g^{s_{att}} | att \in U_k\}$  and  $\text{ACPK}_k = \{V_{k,d} = g^{v_{k,d}} | d \in \mathbb{D}\}$ .

$\mathcal{A}$  gives  $\mathcal{B}$  the  $d^* \in \mathbb{D}$  and  $\mathbb{K}_c \subset \mathbb{K}$  where  $\mathbb{K} \setminus \mathbb{K}_c \neq \emptyset$ . Let  $\mathbb{D}_c = \mathbb{D} \setminus \{d^*\}$ .

$\mathcal{B}$  gives  $\mathcal{A}$   $\text{CMSK}_d = (\alpha_d, \text{SignKey}_d)$  for each  $d \in \mathbb{D}_c$ , and  $\text{AMSK}_k = (\{s_{att} | att \in U_k\}, \{v_{k,d} | d \in \mathbb{D}\})$  for each  $k \in \mathbb{K}_c$ .

$\mathcal{B}$  responds  $\mathcal{A}$ 's key queries as follows:

- $\mathcal{A}$  makes  $\text{CKQ}(gid, d^*)$ .  $\mathcal{B}$  runs the normal  $\text{CKeyGen}$  algorithm, since it knows  $\text{CMSK}_{d^*}$ .
- $\mathcal{A}$  makes  $\text{AKQ}(att, k, \{\text{ucpk}_{gid,d} | d \in \mathbb{D}\})$  where  $(k \in \mathbb{K} \setminus \mathbb{K}_c)$ .  $\mathcal{B}$  runs the normal  $\text{AKeyGen}$  algorithm, since it knows  $\text{AMSK}_k$  for each  $k \in \mathbb{K} \setminus \mathbb{K}_c$ .

$\mathcal{A}$  sends two messages  $M_0$  and  $M_1$  and an LSSS matrix  $(A^*, \rho)$  to  $\mathcal{B}$ . To make the challenge ciphertext,  $\mathcal{B}$  will implicitly set  $g^s$  to be the  $G_{p_1}$  part of  $T$  (we mean that  $T$  is the product of  $g^s \in G_{p_1}$  and possibly an element of  $G_{p_2}$ ).  $\mathcal{B}$  randomly chooses  $v'_2, \dots, v'_n \in \mathbb{Z}_N$  to define a vector  $\vec{v}' = (1, v'_2, \dots, v'_n) \in \mathbb{Z}_N^n$ , and for each row  $x$  of  $A^*$  it randomly picks  $r'_x \in \mathbb{Z}_N$ . Then it chooses a random  $\beta \in \{0, 1\}$  and sets

$$C = M_\beta \cdot \prod_{d=1}^D e(g^{\alpha_d}, T), \quad C' = T,$$

$$\{C_x = T^{aA_x^* \cdot \vec{v}'} T^{-r'_x \cdot s_{\rho(x)}}, \quad C'_x = T^{r'_x} \mid x \in \{1, 2, \dots, l\}\}$$

Note that this implicitly sets  $\vec{v} = s\vec{v}'$  and  $r_x = sr'_x$ . Modulo  $p_1$ , this  $\vec{v}$  is a random vector with first coordinate  $s$  and  $r_x$  is a random value.

If  $T \in G_{p_1}$  (i.e.,  $T = g^s$ ), this is a properly distributed normal ciphertext.

If  $T \in G_{p_1 p_2}$ , we let  $g_2^c$  denote the  $G_{p_2}$  part of  $T$  (i.e.,  $T = g^s g_2^c$ ). We then have a semi-functional ciphertext with  $\vec{u} = ca\vec{v}', \gamma_x = -cr'_x$ , and  $z_{\rho(x)} = s_{\rho(x)}$ . While the values of  $a, v'_2, \dots, v'_n, r'_x, s_{\rho(x)}$  modulo  $p_2$  are uncorrelated from their values modulo  $p_1$  by the Chinese Remainder Theorem, this is a properly distributed semi-functional ciphertext.

Hence, neglecting the probability that  $\Sigma_{sign}$  is broken,  $\mathcal{B}$  can use the output of  $\mathcal{A}$  to break Assumption 1 with advantage  $\epsilon$ . □

**Proof of Lemma 3.4.6.**

*Proof.*  $\mathcal{B}$  is given  $\mathbb{G} = (N, G, G_T, e), g, X_1 X_2, X_3, Y_2 Y_3, T$ . It will simulate **Game** <sub>$j-1,2$</sub>  or **Game** <sub>$j,1$</sub>  to  $\mathcal{A}$ .

$\mathcal{B}$  randomly chooses  $a \in \mathbb{Z}_N$ , then gives  $\mathcal{A}$   $\text{GPK} = \{N, g, h = g^a, X_3\}$ ,  $\mathbb{D} = \{1, 2, \dots, D\}$ ,  $\mathbb{K} = \{1, 2, \dots, K\}$ , the descriptions of  $U_1, U_2, \dots, U_K$ , and the description of secure signature scheme  $\Sigma_{\text{sign}}$ .

For  $d = 1$  to  $D$ ,  $\mathcal{B}$  generates sign key pair  $(\text{SignKey}_d, \text{VerifyKey}_d)$  and randomly chooses  $\alpha_d \in \mathbb{Z}_N$ , then gives  $\mathcal{A}$   $\text{CPK}_d = e(g, g)^{\alpha_d}$  and  $\text{CAPK}_d = \text{VerifyKey}_d$ .

For  $k = 1$  to  $K$ ,  $\mathcal{B}$  randomly chooses  $\{s_{att} \in \mathbb{Z}_N \mid att \in U_k\}$  and  $\{v_{k,d} \in \mathbb{Z}_N \mid d \in \mathbb{D}\}$ , then gives  $\mathcal{A}$   $\text{APK}_k = \{T_{att} = g^{s_{att}} \mid att \in U_k\}$  and  $\text{ACPK}_k = \{V_{k,d} = g^{v_{k,d}} \mid d \in \mathbb{D}\}$ .  $\mathcal{A}$  gives  $\mathcal{B}$  the  $d^* \in \mathbb{D}$  and  $\mathbb{K}_c \subset \mathbb{K}$  where  $\mathbb{K} \setminus \mathbb{K}_c \neq \emptyset$ . Let  $\mathbb{D}_c = \mathbb{D} \setminus \{d^*\}$ .

$\mathcal{B}$  gives  $\mathcal{A}$   $\text{CMSK}_d = (\alpha_d, \text{SignKey}_d)$  for each  $d \in \mathbb{D}_c$ , and  $\text{AMSK}_k = (\{s_{att} \mid att \in U_k\}, \{v_{k,d} \mid d \in \mathbb{D}\})$  for each  $k \in \mathbb{K}_c$ .

$\mathcal{B}$  responds  $\mathcal{A}$ 's key queries as follows:

- $\mathcal{A}$  makes  $\text{CKQ}(gid, d^*)$ . Assume it is the  $i^{\text{th}}$  CKQ and the corresponding user-key will be  $\text{usk}_{gid, d^*}$ ,

- If  $i < j$ ,  $\mathcal{B}$  randomly chooses  $r_{gid, d^*} \in \mathbb{Z}_N$ ,  $R'_{gid, d^*} \in G_{p_3}$ ,  $\{R_{gid, d^*, k} \in G_{p_3} \mid k \in \mathbb{K}\}$ , and sets

$$\text{ucsk}_{gid, d^*} = g^{\alpha_{d^*}} h^{r_{gid, d^*}} (Y_2 Y_3)^{r_{gid, d^*}}, \quad L_{gid, d^*} = g^{r_{gid, d^*}} R'_{gid, d^*},$$

$$\Gamma_{gid, d^*, k} = V_{k, d^*}^{r_{gid, d^*}} R_{gid, d^*, k} (k = 1, 2, \dots, K),$$

$$\sigma_{gid, d^*} = \text{Sign}(\text{SignKey}_{d^*}, gid \parallel d^* \parallel L_{gid, d^*} \parallel \Gamma_{gid, d^*, 1} \parallel \dots \parallel \Gamma_{gid, d^*, K}).$$

(3.8)

Let  $\text{ucpk}_{gid, d^*} = (gid, d^*, L_{gid, d^*}, \{\Gamma_{gid, d^*, k} \mid k \in \mathbb{K}\}, \sigma_{gid, d^*})$ ,

$\mathcal{B}$  answers  $\mathcal{A}$  with  $(\text{ucsk}_{gid, d^*}, \text{ucpk}_{gid, d^*})$ . Note that the values

of  $r_{gid, d^*}$  modulo  $p_2$  and  $p_3$  are uncorrelated to its value modulo

$p_1$ ,  $(\text{ucsk}_{gid,d^*}, \text{ucpk}_{gid,d^*})$  is properly distributed semi-functional user-central-key of type 2.

- If  $i > j$ ,  $\mathcal{B}$  runs the normal CKeyGen algorithm, since it knows  $\text{CMSK}_{d^*}$ .
- If  $i = j$ ,  $\mathcal{B}$  implicitly sets  $g^{r_{gid,d^*}}$  equal to the  $G_{p_1}$  part of  $T$ .  $\mathcal{B}$  randomly chooses  $R_{gid,d^*}, R'_{gid,d^*} \in G_{p_3}, \{R_{gid,d^*,k} \in G_{p_3} | k \in \mathbb{K}\}$ , and sets

$$\begin{aligned} \text{ucsk}_{gid,d^*} &= g^{\alpha_{d^*}} T^a R_{gid,d^*}, \quad L_{gid,d^*} = T R'_{gid,d^*}, \\ \Gamma_{gid,d^*,k} &= T^{v_{k,d^*}} R_{gid,d^*,k} (k = 1, 2, \dots, K), \\ \sigma_{gid,d^*} &= \text{Sign}(\text{SignKey}_{d^*}, gid || d^* || L_{gid,d^*} || \Gamma_{gid,d^*,1} || \dots || \Gamma_{gid,d^*,K}). \end{aligned} \quad (3.9)$$

Let  $\text{ucpk}_{gid,d^*} = (gid, d^*, L_{gid,d^*}, \{\Gamma_{gid,d^*,k} | k \in \mathbb{K}\}, \sigma_{gid,d^*})$ ,  $\mathcal{B}$  answers  $\mathcal{A}$  with  $(\text{ucsk}_{gid,d^*}, \text{ucpk}_{gid,d^*})$ .

If  $T \in G_{p_1 p_3}$ ,  $(\text{ucsk}_{gid,d^*}, \text{ucpk}_{gid,d^*})$  is properly distributed normal user-central-key.

If  $T \in G$ ,  $(\text{ucsk}_{gid,d^*}, \text{ucpk}_{gid,d^*})$  is properly distributed semi-functional user-central-key of type 1. In this case, we have implicitly set  $w_{k,d^*} = v_{k,d^*}$ . If we let  $g_2^b$  denote the  $G_{p_2}$  part of  $T$ , we have that  $\delta = ba$  modulo  $p_2$ , i.e., the  $G_{p_2}$  part of  $\text{ucsk}_{gid,d^*}$  is  $g_2^{ba}$ , the  $G_{p_2}$  part of  $L_{gid,d^*}$  is  $g_2^b$ , and the  $G_{p_2}$  part of  $\Gamma_{gid,d^*,k}$  is  $g_2^{b v_{k,d^*}}$ . Note that the value of  $v_{k,d^*}$  modulo  $p_2$  is uncorrelated from the value of  $v_{k,d^*}$  modulo  $p_1$ .

- $\mathcal{A}$  makes  $\text{AKQ}(att, k, \{\text{ucpk}_{gid,d} | d \in \mathbb{D}\})$  where  $(k \in \mathbb{K} \setminus \mathbb{K}_c)$ .  $\mathcal{B}$  responds  $\mathcal{A}$  as in the real security game by running the normal AKeyGen algorithm, since it knows  $\text{AMSK}_k$  for each  $k \in \mathbb{K} \setminus \mathbb{K}_c$ .

On the  $\text{uask}_{att,gid,d^*}$ , the corresponding  $\text{ucpk}_{gid,d^*}$  must be the answer that  $\mathcal{A}$  gets from  $\mathcal{B}$  by making CKQ because of the signature verification (3.1). Assume the corresponding CKQ is  $i^{\text{th}}$  CKQ, we have

- If  $i < j$ ,

$$\begin{aligned}
 \text{uask}_{att,gid,d^*} &= (\Gamma_{gid,d^*,k})^{s_{att}/v_{k,d^*}} R'_{att,gid,d^*} \\
 &= (V_{k,d^*}^{r_{gid,d^*}} R_{gid,d^*,k})^{s_{att}/v_{k,d^*}} R'_{att,gid,d^*} \\
 &\quad \text{By } \Gamma_{gid,d^*,k} \text{ in (3.8)} \\
 &= T_{att}^{r_{gid,d^*}} R_{att,gid,d^*}.
 \end{aligned}$$

It is properly distributed semi-functional of type 2. Note the corresponding  $(\text{ucsk}_{gid,d^*}, \text{ucpk}_{gid,d^*})$  is properly distributed semi-functional of type 2, we have  $\text{usk}_{gid,d^*}$  as a semi-functional user-key of type 2.

- If  $i > j$ , note that the corresponding  $(\text{ucsk}_{gid,d^*}, \text{ucpk}_{gid,d^*})$  is generated by running the normal CKeyGen algorithm, we have  $\text{usk}_{gid,d^*}$  as a normal user-key.
- If  $i = j$ ,

$$\begin{aligned}
 \text{uask}_{att,gid,d^*} &= (\Gamma_{gid,d^*,k})^{s_{att}/v_{k,d^*}} R'_{att,gid,d^*} \\
 &= (T^{v_{k,d^*}} R_{gid,d^*,k})^{s_{att}/v_{k,d^*}} R'_{att,gid,d^*} \\
 &\quad \text{By } \Gamma_{gid,d^*,k} \text{ in (3.9)} \\
 &= T^{s_{att}} R_{att,gid,d^*}.
 \end{aligned}$$

If  $T \in G_{p_1 p_3}$ ,  $\text{uask}_{att,gid,d^*}$  is a properly distributed normal, where  $g^{r_{gid,d^*}}$  equals to the  $G_{p_1}$  part of  $T$ . Note that the corresponding  $(\text{ucsk}_{gid,d^*}, \text{ucpk}_{gid,d^*})$  is properly distributed normal, we have  $\text{usk}_{gid,d^*}$  as a normal user-key.

If  $T \in G$ ,  $\text{uask}_{att,gid,d^*}$  is a properly distributed semi-functional of type 1. In this case, we have implicitly set  $z_{att} = s_{att}$ . The  $G_{p_2}$  part of  $\text{uask}_{att,gid,d^*}$  is  $g_2^{bs_{att}}$ , and the value of  $s_{att}$  modulo  $p_2$  is uncorrelated from the value of  $s_{att}$  modulo  $p_1$ . Note that the corresponding  $(\text{ucsk}_{gid,d^*},$

$\text{ucpk}_{gid,d^*}$ ) is properly distributed semi-functional of type 1, we have  $\text{usk}_{gid,d^*}$  as a semi-functional user-key of type 1.

$\mathcal{A}$  sends two messages  $M_0$  and  $M_1$  and an LSSS matrix  $(A^*, \rho)$  to  $\mathcal{B}$ . To make the semi-functional challenge ciphertext,  $\mathcal{B}$  implicitly sets  $g^s = X_1$  and  $g_2^c = X_2$ . It randomly chooses  $u_2, \dots, u_n \in \mathbb{Z}_N$  to define a vector  $\vec{u}' = (a, u_2, \dots, u_n)$ , and for each row  $x$  of  $A^*$  it randomly picks  $r'_x \in \mathbb{Z}_N$ . Then it chooses a random  $\beta \in \{0, 1\}$  and sets

$$C = M_\beta \cdot \prod_{d=1}^D e(g^{\alpha_d}, X_1 X_2), \quad C'_x = X_1 X_2,$$

$$\{C_x = (X_1 X_2)^{A^*_x \cdot \vec{u}'} (X_1 X_2)^{-r'_x \cdot s_{\rho(x)}}, \quad C'_x = (X_1 X_2)^{r'_x} \mid x \in \{1, 2, \dots, l\}\}.$$

Note that this implicitly sets  $\vec{v} = sa^{-1}\vec{u}'$ ,  $\vec{u} = c\vec{u}'$ ,  $r_x = sr'_x$ , and  $\gamma_x = -cr'_x$ . The values  $z_{\rho(x)} = s_{\rho(x)}$  match those in the  $j^{\text{th}}$   $\text{usk}_{gid,d^*}$  if it is a semi-functional user-key of type 1, as required.

The challenge ciphertext and the  $\text{usk}_{gid,d^*}$  are *almost* properly distributed, except for the fact that the first coordinate of  $\vec{u}$  (which equals  $ca$ ) is correlated with the value of  $a$  modulo  $p_2$  that also appears in  $\text{ucsk}_{gid,d^*}$  if it is semi-functional. In fact, if the  $\text{usk}_{gid,d^*}$  can be used to reconstruct  $e(g, g)^{\alpha_{d^*} s}$ , we would have  $c\delta - bu_1 = cba - bca = 0$  modulo  $p_2$ , so  $\text{usk}_{gid,d^*}$  is either normal or nominally semi-functional. We must argue that this is hidden to  $\mathcal{A}$ , who *cannot* request any user-keys to reconstruct  $e(g, g)^{\alpha_{d^*} s}$  and then decrypt the challenge ciphertext.

To argue that the value being shared in  $G_{p_2}$  in the challenge ciphertext is information theoretically hidden, we appeal to our restriction that attributes are only used once in labeling the rows of the matrix. Since  $S_{gid} \cup (\bigcup_{k \in \mathbb{K}_c} U_k)$  could not satisfy  $(A^*, \rho)$ ,  $R = \text{span}(\{A^*_x \mid \rho(x) \in S_{gid} \cup (\bigcup_{k \in \mathbb{K}_c} U_k)\})$  does not include the vector  $(1, 0, \dots, 0)$ . Then there exists a vector  $\vec{w}$  such that  $w_1 = \vec{w} \cdot (1, 0, \dots, 0) \neq 0$  and  $\forall \vec{v}_r \in R, \vec{w} \cdot \vec{v}_r = 0$ . We can write  $\vec{u} = \theta \vec{w} + \vec{u}''$  where  $\theta \in \mathbb{Z}_N$  and  $\vec{u}'' \neq \vec{w}$  is distributed

uniformly random in  $R$ . We note that  $\vec{u}''$  reveals no information about  $\theta$  and that  $u_1 = \vec{u} \cdot (1, 0, \dots, 0) = \theta w_1 + \vec{u}'' \cdot (1, 0, \dots, 0)$  can not be determined from  $\vec{u}''$  alone. However, the shares corresponding to rows whose attributes are in  $S_{gid} \cup (\bigcup_{k \in \mathbb{K}_c} U_k)$  only reveal information about  $\vec{u}''$  since  $\forall \vec{v}_r \in R, \vec{w} \cdot \vec{v}_r = 0$ .

$\theta \vec{w}$  appears only in the equations of the form  $A_x^* \cdot \vec{u} + \gamma_x z_{\rho(x)}$  where the attribute  $\rho(x) \notin S_{gid} \cup (\bigcup_{k \in \mathbb{K}_c} U_k)$ . While attributes are only used once in labeling the rows of the matrix, as long as each  $\gamma_x$  is not congruent to 0 modulo  $p_2$ , each of these equations introduces a new unknown  $z_{\rho(x)}$  that appears nowhere else, and thus no information about  $\theta$  can be learned by the adversary. The probability that any  $\gamma_x$  is congruent to 0 modulo  $p_2$  is negligible. Thus, the ciphertext and the key  $\text{usk}_{gid,d^*}$  are properly distributed in the adversary's view with probability negligibly close to 1.

Thus, if  $T \in G_{p_1 p_3}$ ,  $\mathcal{B}$  has properly simulated **Game** <sub>$j-1,2$</sub> , and if  $T \in G$  and all the  $\gamma_x = -cr'_x$  values are non-zero modulo  $p_2$ , then  $\mathcal{B}$  has properly simulated **Game** <sub>$j,1$</sub> . Neglecting the probability that  $\Sigma_{sign}$  is broken,  $\mathcal{B}$  can, therefore, use the output of  $\mathcal{A}$  to gain advantage negligibly close to  $\epsilon$  against Assumption 2.  $\square$

### Proof of Lemma 3.4.7.

*Proof.*  $\mathcal{B}$  is given  $\mathbb{G} = (N, G, G_T, e), g, X_1 X_2, X_3, Y_2 Y_3, T$ . It will simulate **Game** <sub>$j,1$</sub>  or **Game** <sub>$j,2$</sub>  to  $\mathcal{A}$ .

$\mathcal{B}$  randomly chooses  $a \in \mathbb{Z}_N$ , then gives  $\mathcal{A}$   $\text{GPK} = \{N, g, h = g^a, X_3\}$ , as well as  $\mathbb{D} = \{1, 2, \dots, D\}$ ,  $\mathbb{K} = \{1, 2, \dots, K\}$ , the descriptions of  $U_1, U_2, \dots, U_K$ , and the description of secure signature scheme  $\Sigma_{sign}$ .

For  $d = 1$  to  $D$ ,  $\mathcal{B}$  generates sign key pair  $(\text{SignKey}_d, \text{VerifyKey}_d)$  and randomly chooses  $\alpha_d \in \mathbb{Z}_N$ , then gives  $\mathcal{A}$   $\text{CPK}_d = e(g, g)^{\alpha_d}$  and  $\text{CAPK}_d = \text{VerifyKey}_d$ .

For  $k = 1$  to  $K$ ,  $\mathcal{B}$  randomly chooses  $\{s_{att} \in \mathbb{Z}_N | att \in U_k\}$  and  $\{v_{k,d} \in \mathbb{Z}_N | d \in \mathbb{D}\}$ , then gives  $\mathcal{A}$   $\text{APK}_k = \{T_{att} = g^{s_{att}} | att \in U_k\}$  and  $\text{ACPK}_k = \{V_{k,d} = g^{v_{k,d}} | d \in \mathbb{D}\}$ .

$\mathcal{A}$  gives  $\mathcal{B}$  the  $d^* \in \mathbb{D}$  and  $\mathbb{K}_c \subset \mathbb{K}$  where  $\mathbb{K} \setminus \mathbb{K}_c \neq \emptyset$ . Let  $\mathbb{D}_c = \mathbb{D} \setminus \{d^*\}$ .



$\mathcal{B}$  gives  $\mathcal{A}$   $\text{CMSK}_d = (\alpha_d, \text{SignKey}_d)$  for each  $d \in \mathbb{D}_c$ , and  $\text{AMSK}_k = (\{s_{att} | att \in U_k\}, \{v_{k,d} | d \in \mathbb{D}\})$  for each  $k \in \mathbb{K}_c$ .

$\mathcal{B}$  responds  $\mathcal{A}$ 's key queries as follows:

- $\mathcal{A}$  makes  $\text{CKQ}(gid, d^*)$ . Assume it is the  $i^{th}$  CKQ and the corresponding user-key will be  $\text{usk}_{gid, d^*}$ ,

- If  $i < j$ ,  $\mathcal{B}$  randomly chooses  $r_{gid, d^*} \in \mathbb{Z}_N$ ,  $R'_{gid, d^*} \in G_{p_3}$ ,  $\{R_{gid, d^*, k} \in G_{p_3} | k \in \mathbb{K}\}$ , and sets

$$\text{ucsk}_{gid, d^*} = g^{\alpha_{d^*}} h^{r_{gid, d^*}} (Y_2 Y_3)^{r_{gid, d^*}}, \quad L_{gid, d^*} = g^{r_{gid, d^*}} R'_{gid, d^*},$$

$$\Gamma_{gid, d^*, k} = V_{k, d^*}^{r_{gid, d^*}} R_{gid, d^*, k} (k = 1, 2, \dots, K),$$

$$\sigma_{gid, d^*} = \text{Sign}(\text{SignKey}_{d^*}, gid || d^* || L_{gid, d^*} || \Gamma_{gid, d^*, 1} || \dots || \Gamma_{gid, d^*, K}).$$

(3.10)

Let  $\text{ucpk}_{gid, d^*} = (gid, d^*, L_{gid, d^*}, \{\Gamma_{gid, d^*, k} | k \in \mathbb{K}\}, \sigma_{gid, d^*})$ ,  $\mathcal{B}$  answers  $\mathcal{A}$  with  $(\text{ucsk}_{gid, d^*}, \text{ucpk}_{gid, d^*})$ . Note that the values of  $r_{gid, d^*}$  modulo  $p_2$  and  $p_3$  are uncorrelated to its value modulo  $p_1$ ,  $(\text{ucsk}_{gid, d^*}, \text{ucpk}_{gid, d^*})$  is properly distributed semi-functional user-central-key of type 2.

- If  $i > j$ ,  $\mathcal{B}$  runs the normal CKeyGen algorithm, since it knows  $\text{CMSK}_{d^*}$ .
- If  $i = j$ ,  $\mathcal{B}$  implicitly sets  $g^{r_{gid, d^*}}$  equal to the  $G_{p_1}$  part of  $T$ .  $\mathcal{B}$  randomly chooses  $R_{gid, d^*}, R'_{gid, d^*} \in G_{p_3}$ ,  $\{R_{gid, d^*, k} \in G_{p_3} | k \in \mathbb{K}\}$ ,  $\eta \in \mathbb{Z}_N$ , and sets

$$\text{ucsk}_{gid, d^*} = g^{\alpha_{d^*}} T^\eta R_{gid, d^*} (Y_2 Y_3)^\eta, \quad L_{gid, d^*} = T R'_{gid, d^*},$$

$$\Gamma_{gid, d^*, k} = T^{v_{k, d^*}} R_{gid, d^*, k} (k = 1, 2, \dots, K),$$

$$\sigma_{gid, d^*} = \text{Sign}(\text{SignKey}_{d^*}, gid || d^* || L_{gid, d^*} || \Gamma_{gid, d^*, 1} || \dots || \Gamma_{gid, d^*, K}).$$

(3.11)

Let  $\text{ucpk}_{gid,d^*} = (gid, d^*, L_{gid,d^*}, \{\Gamma_{gid,d^*,k} \mid k \in \mathbb{K}\}, \sigma_{gid,d^*})$ ,  $\mathcal{B}$  answers  $\mathcal{A}$  with  $(\text{ucsk}_{gid,d^*}, \text{ucpk}_{gid,d^*})$ .

If  $T \in G_{p_1 p_3}$ ,  $(\text{ucsk}_{gid,d^*}, \text{ucpk}_{gid,d^*})$  is properly distributed semi-functional user-central-key of type 2.

If  $T \in G$ ,  $(\text{ucsk}_{gid,d^*}, \text{ucpk}_{gid,d^*})$  is properly distributed semi-functional user-central-key of type 1. In this case, we have implicitly set  $w_{k,d^*} = v_{k,d^*}$ . If we let  $g_2^b$  denote the  $G_{p_2}$  part of  $T$  and  $g_2^{y_2} = Y_2$ , we have that  $\delta = ba + \eta y_2$  modulo  $p_2$ . i.e., the  $G_{p_2}$  part of  $\text{ucsk}_{gid,d^*}$  is  $g_2^{ba+\eta y_2}$ , the  $G_{p_2}$  part of  $L_{gid,d^*}$  is  $g_2^b$ , and the  $G_{p_2}$  part of  $\Gamma_{gid,d^*,k}$  is  $g_2^{bv_{k,d^*}}$ . Note that the value of  $v_{k,d^*}$  modulo  $p_2$  is uncorrelated from the value of  $v_{k,d^*}$  modulo  $p_1$ .

- $\mathcal{A}$  makes  $\text{AKQ}(att, k, \{\text{ucpk}_{gid,d} \mid d \in \mathbb{D}\})$  where  $(k \in \mathbb{K} \setminus \mathbb{K}_c)$ .  $\mathcal{B}$  responds  $\mathcal{A}$  as in the real security game by running the normal **AKeyGen** algorithm, since it knows  $\text{AMSK}_k$  for each  $k \in \mathbb{K} \setminus \mathbb{K}_c$ .

On the  $\text{uask}_{att,gid,d^*}$ , the corresponding  $\text{ucpk}_{gid,d^*}$  must be the answer that  $\mathcal{A}$  gets from  $\mathcal{B}$  by making CKQ because of the signature verification (3.1). Assume the corresponding CKQ is  $i^{th}$  CKQ, we have

- If  $i < j$ ,

$$\begin{aligned} \text{uask}_{att,gid,d^*} &= (\Gamma_{gid,d^*,k})^{s_{att}/v_{k,d^*}} R'_{att,gid,d^*} \\ &= (V_{k,d^*}^{r_{gid,d^*}} R_{gid,d^*,k})^{s_{att}/v_{k,d^*}} R'_{att,gid,d^*} \\ &\quad \text{By } \Gamma_{gid,d^*,k} \text{ in (3.10)} \\ &= T_{att}^{r_{gid,d^*}} R_{att,gid,d^*}. \end{aligned}$$

It is properly distributed semi-functional of type 2.

Note the corresponding  $(\text{ucsk}_{gid,d^*}, \text{ucpk}_{gid,d^*})$  is properly distributed semi-functional of type 2, we have  $\text{usk}_{gid,d^*}$  as a semi-functional user-key

of type 2.

- If  $i > j$ , note that the corresponding  $(\text{ucsk}_{gid,d^*}, \text{ucpk}_{gid,d^*})$  is generated by running the normal CKeyGen algorithm, we have  $\text{usk}_{gid,d^*}$  as a normal user-key.
- If  $i = j$ ,

$$\begin{aligned}
 \text{uask}_{att,gid,d^*} &= (\Gamma_{gid,d^*,k})^{s_{att}/v_{k,d^*}} R'_{att,gid,d^*} \\
 &= (T^{v_{k,d^*}} R_{gid,d^*,k})^{s_{att}/v_{k,d^*}} R'_{att,gid,d^*} \\
 &\quad \text{By } \Gamma_{gid,d^*,k} \text{ in (3.11)} \\
 &= T^{s_{att}} R_{att,gid,d^*}.
 \end{aligned}$$

If  $T \in G_{p_1 p_3}$ ,  $\text{uask}_{att,gid,d^*}$  is properly distributed semi-functional of type 2, where  $g^{r_{gid,d^*}}$  equals to the  $G_{p_1}$  part of  $T$ . Note that the corresponding  $(\text{ucsk}_{gid,d^*}, \text{ucpk}_{gid,d^*})$  is properly distributed semi-functional of type 2, we have  $\text{usk}_{gid,d^*}$  as a semi-functional user-key of type 2.

If  $T \in G$ ,  $\text{uask}_{att,gid,d^*}$  is properly distributed semi-functional of type 1. In this case, we have implicitly set  $z_{att} = s_{att}$ . The  $G_{p_2}$  part of  $\text{uask}_{att,gid,d^*}$  is  $g_2^{b s_{att}}$ , and the value of  $s_{att}$  modulo  $p_2$  is uncorrelated from the value of  $s_{att}$  modulo  $p_1$ . Note that the corresponding  $(\text{ucsk}_{gid,d^*}, \text{ucpk}_{gid,d^*})$  is properly distributed semi-functional of type 1, we have  $\text{usk}_{gid,d^*}$  as a semi-functional user-key of type 1.

$\mathcal{A}$  sends two messages  $M_0$  and  $M_1$  and an LSSS matrix  $(A^*, \rho)$  to  $\mathcal{B}$ . To make the semi-functional challenge ciphertext,  $\mathcal{B}$  implicitly sets  $g^s = X_1$  and  $g_2^c = X_2$ . It randomly chooses  $u_2, \dots, u_n \in \mathbb{Z}_N$  to define a vector  $\vec{u}' = (a, u_2, \dots, u_n)$ , and for each row  $x$  of  $A^*$  it randomly picks  $r'_x \in \mathbb{Z}_N$ . Then it chooses a random  $\beta \in \{0, 1\}$  and

sets

$$C = M_\beta \cdot \prod_{d=1}^D e(g^{\alpha_d}, X_1 X_2), \quad C' = X_1 X_2,$$

$$\{C_x = (X_1 X_2)^{A_x^* \cdot \vec{u}'} (X_1 X_2)^{-r'_x \cdot s_{\rho(x)}}, \quad C'_x = (X_1 X_2)^{r'_x} \mid x \in \{1, 2, \dots, l\}\}$$

We note that this implicitly sets  $\vec{v} = sa^{-1}\vec{u}'$ ,  $\vec{u} = c\vec{u}'$ ,  $r_x = sr'_x$  and  $\gamma_x = -cr'_x$ . The values  $z_{\rho(x)} = s_{\rho(x)}$  match those in the  $j^{\text{th}}$   $\text{usk}_{gid,d^*}$  if it is a semi-functional user-key of type 1, as required.

The challenge ciphertext and the  $\text{usk}_{gid,d^*}$  are properly distributed because the  $G_{p_2}$  part of  $\text{ucsk}_{gid,d^*}$  is randomized by  $\eta$ .

Thus, if  $T \in G_{p_1 p_3}$ ,  $\mathcal{B}$  has properly simulated **Game**<sub>j,2</sub>, and if  $T \in G$ ,  $\mathcal{B}$  has properly simulated **Game**<sub>j,1</sub>. Neglecting the probability that  $\Sigma_{\text{sign}}$  is broken,  $\mathcal{B}$  can, therefore, use the output of  $\mathcal{A}$  to gain advantage  $\epsilon$  against Assumption 2.  $\square$

### Proof of Lemma 3.4.8.

*Proof.*  $\mathcal{B}$  is given  $\mathbb{G} = (N, G, G_T, e), g, g^\alpha X_2, X_3, g^s Y_2, Z_2, T$ . It will simulate **Game**<sub>q,2</sub> or **Game**<sub>Final</sub> to  $\mathcal{A}$ .

$\mathcal{B}$  randomly chooses  $a \in \mathbb{Z}_N$ , then gives  $\mathcal{A}$   $\text{GPK} = \{N, g, h = g^a, X_3\}$ , as well as  $\mathbb{D} = \{1, 2, \dots, D\}$ ,  $\mathbb{K} = \{1, 2, \dots, K\}$ , the descriptions of  $U_1, U_2, \dots, U_K$ , and the description of secure signature scheme  $\Sigma_{\text{sign}}$ .  $\mathcal{B}$  randomly chooses  $d' \in \mathbb{D}$ .

For  $d = 1$  to  $D$ ,  $\mathcal{B}$  generates sign key pair  $(\text{SignKey}_d, \text{VerifyKey}_d)$ . If  $d \neq d'$ ,  $\mathcal{B}$  randomly chooses  $\alpha_d \in \mathbb{Z}_N$  and sets  $\text{CPK}_d = e(g, g)^{\alpha_d}$ , otherwise  $\mathcal{B}$  sets  $\text{CPK}_d = e(g, g^\alpha X_2)$ .  $\mathcal{B}$  gives  $\mathcal{A}$   $\text{CPK}_d$  and  $\text{CAPK}_d = \text{VerifyKey}_d$ . This implicitly sets  $\alpha_{d'} = \alpha$ . For  $k = 1$  to  $K$ ,  $\mathcal{B}$  randomly chooses  $\{s_{att} \in \mathbb{Z}_N \mid att \in U_k\}$  and  $\{v_{k,d} \in \mathbb{Z}_N \mid d \in \mathbb{D}\}$ , then gives  $\mathcal{A}$   $\text{APK}_k = \{T_{att} = g^{s_{att}} \mid att \in U_k\}$  and  $\text{ACPK}_k = \{V_{k,d} = g^{v_{k,d}} \mid d \in \mathbb{D}\}$ .  $\mathcal{A}$  gives  $\mathcal{B}$  the  $d^* \in \mathbb{D}$  and  $\mathbb{K}_c \subset \mathbb{K}$  where  $\mathbb{K} \setminus \mathbb{K}_c \neq \emptyset$ . If  $d^* \neq d'$ , it means  $\mathcal{B}$  guesses a wrong  $d^*$ ,  $\mathcal{B}$  aborts. Note that the public parameters are properly distributed,

the probability  $\mathcal{B}$  guesses the right  $b^*$  is  $\frac{1}{D}$ .

$\mathcal{B}$  gives  $\mathcal{A}$   $\text{CMSK}_d = (\alpha_d, \text{SignKey}_d)$  for each  $d \in \mathbb{D}_c = \mathbb{D} \setminus \{d^*\}$ , and  $\text{AMSK}_k = (\{s_{att} | att \in U_k\}, \{v_{k,d} | d \in \mathbb{D}\})$  for each  $k \in \mathbb{K}_c$ .

$\mathcal{B}$  responds  $\mathcal{A}$ 's key queries as follows:

- $\mathcal{A}$  makes  $\text{CKQ}(gid, d^*)$ . Assume the corresponding user-key will be  $\text{usk}_{gid, d^*}$ ,  $\mathcal{B}$  randomly chooses  $r_{gid, d^*} \in \mathbb{Z}_N, R_{gid, d^*}, R'_{gid, d^*} \in G_{p_3}, \{R_{gid, d^*, k} \in G_{p_3} | k \in \mathbb{K}\}$ , and sets

$$\begin{aligned} \text{ucsk}_{gid, d^*} &= (g^\alpha X_2) h^{r_{gid, d^*}} (Z_2)^{r_{gid, d^*}} R_{gid, d^*}, \quad L_{gid, d^*} = g^{r_{gid, d^*}} R'_{gid, d^*}, \\ \Gamma_{gid, d^*, k} &= V_{k, d^*}^{r_{gid, d^*}} R_{gid, d^*, k} (k = 1, 2, \dots, K), \\ \sigma_{gid, d^*} &= \text{Sign}(\text{SignKey}_{d^*}, gid || d^* || L_{gid, d^*} || \Gamma_{gid, d^*, 1} || \dots || \Gamma_{gid, d^*, K}). \end{aligned} \quad (3.12)$$

Let  $\text{ucpk}_{gid, d^*} = (gid, d^*, L_{gid, d^*}, \{\Gamma_{gid, d^*, k} | k \in \mathbb{K}\}, \sigma_{gid, d^*})$ ,  $\mathcal{B}$  answers  $\mathcal{A}$  with  $(\text{ucsk}_{gid, d^*}, \text{ucpk}_{gid, d^*})$ . Note that the value of  $r_{gid, d^*}$  modulo  $p_2$  is uncorrelated to its value modulo  $p_1$ ,  $(\text{ucsk}_{gid, d^*}, \text{ucpk}_{gid, d^*})$  is properly distributed semi-functional user-central-key of type 2.

- $\mathcal{A}$  makes  $\text{AKQ}(att, k, \{\text{ucpk}_{gid, d} | d \in \mathbb{D}\})$  where  $(k \in \mathbb{K} \setminus \mathbb{K}_c)$ .  $\mathcal{B}$  responds  $\mathcal{A}$  as in the real security game by running the normal  $\text{AKeyGen}$  algorithm, since it knows  $\text{AMSK}_k$  for each  $k \in \mathbb{K} \setminus \mathbb{K}_c$ .

On the  $\text{uask}_{att, gid, d^*}$ , the corresponding  $\text{ucpk}_{gid, d^*}$  must be the answer that  $\mathcal{A}$  gets from  $\mathcal{B}$  by making  $\text{CKQ}$  because of the signature verification (3.1). Then we have

$$\begin{aligned} \text{uask}_{att, gid, d^*} &= (\Gamma_{gid, d^*, k})^{s_{att}/v_{k, d^*}} R'_{att, gid, d^*} \\ &= (V_{k, d^*}^{r_{gid, d^*}} R_{gid, d^*, k})^{s_{att}/v_{k, d^*}} R'_{att, gid, d^*} \quad \text{By } \Gamma_{gid, d^*} \text{ in (3.12)} \\ &= T_{att}^{r_{gid, d^*}} R_{att, gid, d^*}. \end{aligned}$$

It is properly distributed semi-functional of type 2. Note that the corresponding  $(\text{ucsk}_{gid,d^*}, \text{ucpk}_{gid,d^*})$  is properly distributed semi-functional of type 2, we have  $\text{usk}_{gid,d^*}$  as a semi-functional user-key of type 2.

$\mathcal{A}$  sends two messages  $M_0, M_1$  and an LSSS matrix  $(A^*, \rho)$  to  $\mathcal{B}$ . To make the semi-functional challenge ciphertext,  $\mathcal{B}$  implicitly takes  $s$  from the assumption term  $g^s Y_2$  and sets  $g_2^c = Y_2$ . It randomly chooses  $u_2, \dots, u_n \in \mathbb{Z}_N$  to define a vector  $\vec{u}' = (a, u_2, \dots, u_n)$ , and for each row  $x$  of  $A^*$  it randomly picks  $r'_x \in \mathbb{Z}_N$ . Then it chooses a random  $\beta \in \{0, 1\}$  and sets

$$C = M_\beta \cdot T \cdot \prod_{d \in \mathbb{D}, d \neq d^*} e(g, g)^{\alpha_{d^*} s}, \quad C' = g^s Y_2,$$

$$\{C_x = (g^s Y_2)^{A_x^* \cdot \vec{u}'} (g^s Y_2)^{-r'_x \cdot s_{\rho(x)}}, \quad C'_x = (g^s Y_2)^{r'_x} \mid x \in \{1, 2, \dots, l\}\}.$$

We note that this implicitly sets  $\vec{v} = sa^{-1} \vec{u}'$ ,  $\vec{u} = cu'$ ,  $r_x = sr'_x$  and  $\gamma_x = -cr'_x$ .

If  $T = e(g, g)^{\alpha s}$ , this is a properly distributed semi-functional ciphertext of the encryption of  $M_\beta$ . Otherwise, it is a properly distributed semi-functional ciphertext of the encryption of a random message in  $G_T$ . Thus, neglecting the probability that  $\Sigma_{sign}$  is broken,  $\mathcal{B}$  can use the output of  $\mathcal{A}$  to gain advantage  $\frac{\epsilon}{D}$  against Assumption 3.  $\square$

### 3.5 Interval Encryption Schemes

A broadcast encryption (BE) scheme enables a broadcaster to choose a subset  $S$  of  $n$  users, who are listening to the broadcast channel, and encrypt a message for this subset. Any user in  $S$  is allowed to successfully decrypt the message. Even if all the users outside of  $S$  collude together, they can not obtain any useful information of the broadcast message. In the following, we also use  $r$  to represent the number of revoked users, i.e.,  $r = n - |S|$  where  $|S|$  is the size of  $S$ . Compared with a private key BE scheme [13, 288], a public key broadcast encryption has the benefit that users are not required to pre-share any private information. Therefore, in this section, we mainly

focus on public key broadcast encryption. Three efficiency parameters of a BE scheme are of our major concern: the transmission cost, user storage, and the decryption time.

The transmission cost of most current public key BE constructions will grow with the increase of the revocation number  $r$ . Naor et al. [221] presented a BE construction (NNL method) with an average ciphertext size of  $1.38r$  and private key size  $\mathcal{O}(\log^2 n)$ . The private key size is further improved to  $\mathcal{O}(\log^{1+\epsilon} n)$ ,  $0 < \epsilon < 1$  in the HS construction [134], where the ciphertext size blows up with a  $\frac{1}{\epsilon}$  factor. The private key size is further improved to  $\mathcal{O}(\log n)$  by Goodrich et al. [126]. Dodis and Fazio [101] presented a generic method (DF transformation) to transform the NNL method and HS construction into a public key broadcast system using hierarchical identity-based encryption (HIBE). The transmission overhead remains unchanged and the private key consists of  $\mathcal{O}(\log^2 n)$  and  $\mathcal{O}(\log^{1+\epsilon} n)$  HIBE node secret keys if DF transformation is instantiated with BBG HIBE [35]. The security is reduced to standard Decisional BDHE assumption and the decryption time cost is  $\mathcal{O}(\log n)$ . The decryption time is then improved to constant by Liu and Teng [200]. However, their security is reduced to decisional BDH assumption in the random oracle model. Recently, Sahai and Waters proposed a BE system with a transmission cost linearly dependent on  $r$  and constant storage cost. However, the decryption cost is linearly dependent on  $r$  and the security is reduced to a complex assumption called  $q$ -MEBDH assumption. Actually, it has been pointed out [159] that at least a single key per revoked user should be included in the transmission cost and hence,  $r$  might be the lower bound of the transmission overhead in any BE scheme with reasonable decryption computational and storage cost. Therefore, constructing a BE system with a transmission overhead lower than  $r$  as well as reasonable user storage and computational cost is still an open problem, which is one of the major issues of this section.

On the other hand, there are two major application scenarios [39] for BE: applications where we broadcast large sets, namely sets of size  $n-r$  for  $r \ll n$  and applications

where we broadcast small sets, namely sets of size  $|S| \ll n$ . Apparently, a BE system with a transmission cost dependent on  $r$  is not efficient when  $r$  grows, and especially it fails to be an optimal choice for the second kind of application where  $r$  is very close to  $n$ . Before BGW proposed their construction [39], the only suitable solution for the latter scenario is the trivial solution, i.e., encrypting the message under each recipient's key.

In order to construct a BE scheme suitable for *arbitrary* receiver sets, we need to break the barrier of  $r$ . BGW [39] proposed an elegant BE scheme with constant size ciphertext as the first attempt to solve this problem. Although the ciphertext and private key size of their construction is constant, the size of public keys is linearly dependent on  $n$ . The public key must be accessible to any decryptor, which implies a high storage cost of size  $\mathcal{O}(n)$ . This makes their system unsuitable for the application scenario where users have only limited storage capability [240]. Their underlying assumption is the standard Decisional BDHE assumption. Later, Deleralee [95] proposed a BE construction where the public key size depends on the maximum size of  $S$  while both ciphertext and private key remain constant size. However, this still does not serve as an efficient solution for applications where the receiver set is large, namely  $r \ll n$ . The security of this construction is reduced to a complex assumption called GDHE assumption in the random oracle model. Besides, the decryption of both constructions is not efficient. The decryption cost of the BGW construction depends on  $n$ , and the decryption of Deleralee's construction requires  $\mathcal{O}(|S|)$  operations.

In this section, we consider this problem from a brand-new angle and a more practical point of view. The basic motivation comes from the following observation: in a BE system with  $n$  users, where each user is assigned with an index  $i \in [1, n]$ . The receiver set  $S$  can be regarded as a collection of  $k$  intervals. Considering the fact that the number of intervals containing in  $S$  is always less than  $r + 1$  and in the best cases  $k$  could even be much less than  $r$ , the system performance can be dramatically increased if the



transmission overhead of the BE system is only determined by the interval number  $k$  while irrelevant of  $r$ . In this study, we will use more detailed performance analysis and simulation to show that a BE construction based on  $k$  is always more efficient than the previous scheme dependent on  $r$ , and suitable for more cases in practice.

In order to realize a BE system with a transmission overload dependent on  $k$ , we propose a new type of encryption called interval encryption. In interval encryption, a message is encrypted under a collection of natural intervals  $S = \bigcup_{j=1}^k NI_j$ , where  $NI_j$  is a natural interval in  $[1, n]$ . Each receiver is identified by a unique natural number  $i \in [1, n]$  and assigned with the respective private key. The decryption is successful if and only if the natural number  $i$  belongs to  $S$ .

We present a generic methodology which can transform a series of binary tree encryptions into interval encryptions. We illustrate the basic methodology using the BBG HIBE scheme [35]. The construction achieves a ciphertext size of  $\mathcal{O}(k)$ , and  $\mathcal{O}(\log n)$  private storage. The decryption is dominated by at most  $\mathcal{O}(\log n)$  group operations. The security is reduced to the Decisional BDHE assumption. We note that one of the best public key BE schemes under this assumption is the DF transformation of the HS construction which requires a transmission cost of  $\mathcal{O}(r/\epsilon)$  size and the private key consists of  $\mathcal{O}(\log^{1+\epsilon} n)$  HIBE node secret keys, where  $0 < \epsilon < 1$ .

We also apply our basic methodology to the fully secure HIBE [182] scheme proposed by Lewko and Waters to present an adaptively secure interval encryption scheme. Gentry and Waters [123] proposed the first adaptively secure BE scheme under a complex bilinear assumption. The public parameter size of their construction is of  $\mathcal{O}(|S|)$ . The private key size is constant, and the ciphertext size of their construction is of  $\mathcal{O}(\max |S|)$ . Later, Waters [290] gave the first short ciphertext adaptively secure BE system under static (i.e. non  $q$ -based) assumptions. However, both of the public parameter and private key size are linearly dependent on  $n$ . The public parameter of our construction is of size  $\mathcal{O}(\log n)$  and the ciphertext size is of  $\mathcal{O}(k)$ . It only requires

$\mathcal{O}(\log n)$  private storage. In other words, our construction serves as one of the most efficient adaptively secure BE systems. Besides, our construction also reduces its security to static assumptions.

Since we consider the proposal of this new concept and the corresponding methodology of one of our major contributions, an inclusive extended interval encryption is proposed as another illustration of the power of our basic methodology. A message is encrypted under a collection of intervals  $S = \bigcup_{j=1}^k NI_j$  in this extended construction. A user's private key corresponds to a certain interval  $NI_\omega$ . The decryption is successful if and only if there's at least one interval  $NI_j, j \in [1, k]$  such that  $NI_\omega \subseteq NI_j$ . The construction also provides user with delegation capability. We also discuss several interesting applications of interval encryption. In particular, we propose a useful concept of range attribute-based encryption and present an efficient construction from interval encryption.

### 3.5.1 Definitions

**Assumptions.** Bilinear maps [223] are crucial for our construction. A pairing is an efficiently computable, nondegenerate function,  $\hat{e} : \mathbb{G}_1 \times \mathbb{G}_1 \rightarrow \mathbb{G}_2$ , with the bilinearity property that  $\hat{e}(g^r, g^s) = \hat{e}(g, g)^{rs}$ . Here,  $\mathbb{G}_1$ , and  $\mathbb{G}_2$  are all multiplicative groups of prime order  $p$ , respectively, generated by  $g$  and  $\hat{e}(g, g)$ .

The security proof of our constructions relies on the Decisional  $d + 1$  BDHE assumption, which can be stated as [47]: Given a tuple  $[h, g, g^\alpha, g^{(\alpha^2)}, \dots, g^{\alpha^d}, g^{(\alpha^{d+2})}, \dots, g^{(\alpha^{2d})}, Z] \in \mathbb{G}_1^{2d+1} \times \mathbb{G}_2$  for a random exponent  $\alpha \in \mathbb{Z}_p$ , decide whether  $Z = \hat{e}(g, h)^{\alpha^{d+1}}$ .

**Notations.** We inherit most notations from the underlying BTE and FSE [62] construction. Recall that  $d$  denotes the depth of the tree, and  $n = 2^d$  is the number of leaf nodes. We set the root node to be  $\varepsilon$  by convention. The other nodes on the tree have an associated name chosen from  $\{0, 1\}^{\leq d}$ . The left child of a node is concatenated

with 0, and the right child is concatenated with 1. Therefore, each leaf node will also have an associated binary name  $[\omega_1\omega_2 \cdots \omega_d]$ . We also let a natural number  $\omega \in [1, n]$  associated with the  $\omega$ -th leaf node of the binary tree (starting from left to right). We implicitly let  $\omega = [\omega_1\omega_2 \cdots \omega_d]$  in the remainder of this section. The  $j$ -bit prefix of a string  $\omega = [\omega_1\omega_2 \cdots \omega_d]$  is denoted by  $\omega|_j$ , namely  $\omega|_j = [\omega_1\omega_2 \cdots \omega_j]$ . We implicitly set  $\omega|_0 = \varepsilon$  and  $\omega|_d = \omega$ . It is easy to observe that a set of nodes  $\omega|_j, j \in [1, d]$  corresponds to the nodes on the path from the root to the leaf node  $\omega$  (see Figure 3.11. (b)). Besides, we use  $\omega|_{j,(RS)}$  or  $\omega|_{j,(LS)}$  to denote the right or left sibling of  $\omega|_j$ , respectively, if  $\omega|_j$  has such a sibling. Namely,  $\omega|_{j,(RS)} = [\omega_1\omega_2 \cdots \omega_{j-1}1]$  or  $\omega|_{j,(LS)} = [\omega_1\omega_2 \cdots \omega_{j-1}0]$ .

Generally, our BE system consists of two parallel BTE systems: the right BTE system and the left BTE system. The right BTE system covers all the leaf nodes in the interval  $[\omega, n]$  and the left BTE system covers all the leaf nodes in the interval  $[1, \omega]$ . User  $\omega$  is assigned with a unique right master key and a left master key. All the node secret keys or private keys for  $\omega$  in the right BTE system are derived from the right master key and in the left BTE system are derived from the left master key. We use two different subindexes (L) or (R) in the notations of all these keys to distinguish the left or right BTE system they correspond to respectively.

### 3.5.2 Security Models

Our construction is a key encapsulation mechanism (KEM)<sup>7</sup>, thus long messages can be encrypted under a short symmetric key. An interval encryption scheme is made up of four randomized algorithms:

**Setup**( $n$ ): Take as input a natural interval  $[1, n]$ . It outputs a public key  $PK$  and the system master key  $SK_\varepsilon$ .

**PvkGen**( $\omega, SK_\varepsilon$ ): Take as input a natural number  $\omega \in [1, n]$  and the system master

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<sup>7</sup>We adopt KEM for the ease of comparison since all the BE constructions in the literature employ the same mechanism.

key  $SK_\varepsilon$ . It outputs a private key  $D_\omega$ .

**Encrypt**( $S, PK$ ): Take as input a public key  $PK$ , and a  $k$ -wise natural interval set  $S = \bigcup_{j=1}^k NI_j$  where  $NI_j = [l_j, r_j]$  satisfying  $1 \leq l_1 \leq r_1 < l_2 \leq r_2 \cdots < l_k \leq r_k \leq n$ . For  $j \in [1, k]$ , it outputs  $k$  pairs  $\{Hdr_j, K_j\}$ . We call  $Hdr = \{Hdr_j\}_{j=1}^k$  the header and  $K = \{K_j\}_{j=1}^k$  the message encryption keys.

Let  $M$  be a message that should be decipherable precisely by the receivers holding the private key corresponding to  $\omega \in S$ . For  $j \in [1, k]$ , let  $C_j$  be the encryption of  $M$  under the message encryption key  $K_j$ . Let  $C_M$  be the collection of these encryption, namely  $C_M = \{C_j\}_{j=1}^k$ . The whole ciphertext consists of  $(S, Hdr, C_M)$ .

**Decrypt** ( $S, \omega, D_\omega, Hdr, PK$ ): Take as input a  $k$ -wise natural interval set  $S = \bigcup_{j=1}^k NI_j$  and the private key  $D_\omega$  for a natural number  $\omega \in [1, n]$ , a header  $Hdr$ , a public key  $PK$ . If  $\omega \in NI_j, 1 \leq j \leq k$ , the algorithm outputs the corresponding message encryption key  $K_j \in \mathcal{K}$ .

The system is considered correct, if for all  $k$ -wise natural interval sets  $S = \bigcup_{j=1}^k NI_j$  and natural numbers  $\omega \in NI_j$  (where  $j \in [1, k]$ ), if

$$PK \xleftarrow{R} \mathbf{Setup}(n), D_\omega \xleftarrow{R} \mathbf{PvkGen}(\omega, SK_\varepsilon), (Hdr, K) \xleftarrow{R} \mathbf{Encrypt}(S, PK),$$

then  $\mathbf{Decrypt}(S, \omega, D_\omega, Hdr, PK) = K_j$ . The concept of interval encryption is close to private linear BE (PLBE) mentioned in [42], and can be viewed as an extension of PLBE.

**Semantic Security**(IND-sI-CPA): The selective interval game is very similar to that of BE [95], and it forms as follow:

**Init** The adversary outputs a  $k$ -wise natural interval set  $S^* = \bigcup_{j=1}^k NI_j^*$ , where  $NI_j^* = [l_j^*, r_j^*]$  satisfying  $1 \leq l_1^* \leq r_1^* < l_2^* \leq r_2^* \cdots < l_k^* \leq r_k^* \leq n$ , which it wishes to attack.

**Setup** The challenger runs  $\mathbf{Setup}(n)$  to obtain a public key  $PK$  for the adversary.

**Phase 1** The adversary issues query for private key of  $\omega \notin S^*$ .

**Challenge** The challenger runs algorithm **Encrypt** to obtain  $(\text{Hdr}^*, K) \xleftarrow{R} \text{Encrypt}(S^*, PK)$  where  $K \in \mathcal{K}^k$ . Next, the challenger picks a random  $\beta \in \{0, 1\}$ . It sets  $K^* = K$  if  $\beta = 1$  and sets  $K^*$  to a random string of length equal to  $|K|$  otherwise. It then sends  $\text{Hdr}^*, K^*$  to the adversary.

**Phase 2** Same as phase 1.

**Guess** The adversary outputs its guess  $\beta' \in \{0, 1\}$  for  $\beta$  and wins the game if  $\beta' = \beta$ .

The adversary's advantage is the absolute value of the difference between its success probability and  $\frac{1}{2}$ .

**Definition 3.5.1.** An interval encryption scheme is selective-interval chosen plaintext secure (IND-sI-CPA) if all polynomial time adversaries have at most a negligible advantage in winning the above security game.

The adaptive CPA security can be defined in a similar way except that there is no **Init** stage in the adaptive game and the challenge interval  $S^*$  in the **Challenge** stage should be provided under the restriction that none of the identities  $\omega$  for the key queries of **Phase 1** and **Phase 2** belongs to  $S^*$ , i.e.,  $\omega \notin S^*$ .

The ultimate security goal is to realize IND-CCA security where the adversary doesn't need to choose the interval set at the beginning and is provided with a decryption oracle. However, we only concentrate on IND-sI-CPA security.

### 3.5.3 Binary Tree Encryption and Forward Secure Encryption

The concept of binary tree encryption (BTE) was first proposed by Canetti, et al [64]. BTE is a relaxation of hierarchical identity-based encryption (HIBE) [122]. As in HIBE, a “master” public key  $PK$  is associated with a binary tree in BTE; each node  $\omega$  in this tree has a corresponding secret key  $SK_\omega$ . To encrypt a message “targeted” for some node, one uses both  $PK$  and the name of the target node; the resulting ciphertext can then be decrypted using the secret key of the target node. Moreover, as

in HIBE the secret key of any node can be employed to derive the secret keys for the children of that node. The only difference between HIBE and BTE is that the latter insists on a *binary* tree, where each non-leaf node only has two child nodes.

Technically speaking, forward secure encryption (FSE) is an elegant application of BTE. Let the depth of a binary tree be  $d$  which implies it has  $n = 2^d$  leaf nodes. In a FSE scheme, the lifetime of a system is divided into  $n = 2^d$  time periods, each of which is associated with a unique leaf node of the tree. A user holding a private key for time period  $\omega$  can open all the messages encrypted under the subsequent time periods, namely  $\omega' \in [\omega, n]$ . The private key  $D_\omega$  in the FSE construction contains the node secret keys  $SK_\omega$  for the leaf node  $\omega$  as well as node secret keys for the right siblings of the nodes on the path from the root to node  $\omega$ , where all these node secret keys come from the underlying BTE scheme. To encrypt a message for a certain period  $\omega'$ , one uses both  $PK$  and the name of respective leaf node  $\omega'$  as in the BTE scheme; the resulting ciphertext can then be decrypted using node secret key  $SK_{\omega'}$ , which is also similar to the BTE scheme. As shown in Figure 3.11. (a), a private key  $D_2$  containing the node secret keys  $SK_2$ ,  $SK_c$ , and  $SK_b$  can be used to derive all the node secret keys for leaf nodes falling into the interval  $[2, 8]$ . Therefore,  $D_2$  can be used to open all the messages encrypted under time periods in the interval  $[2, 8]$ .

Indeed, FSE can be viewed as a special case of interval encryption. As shown in Figure 3.11. (a), if we use ciphertext  $C_4$  encrypted under leaf node 4 to represent the interval  $[1, 4]$ , then only the private key for time period  $\omega \in [1, 4]$  can be used to open the message, e.g.,  $D_2$  could be used for the decryption of  $C_4$  because  $SK_4$  can be derived from  $SK_c$ , which belongs to  $D_2$ . However,  $D_5$  cannot be used for decrypting  $C_4$  as it is impossible to deduce  $SK_4$  from any node secret keys included in  $D_5$ .

In the remainder of this section, we use a right direction arrow from a certain leaf node (or the corresponding index in the axis) to denote this particular private key distribution mode. A right direction arrow from a leaf node  $\omega$  means that all the node

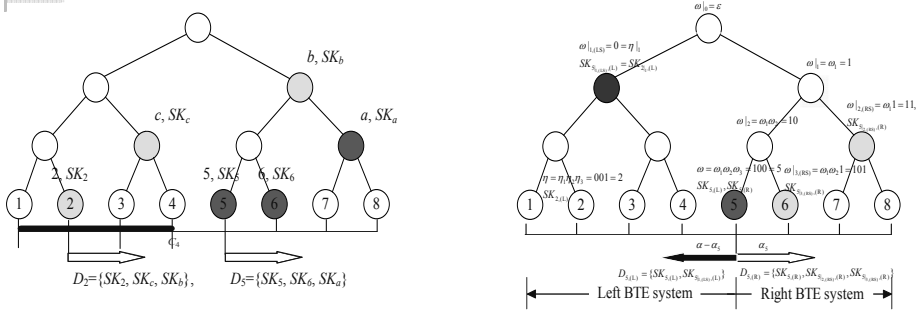


Figure 3.11: (a). The key distribution mode of forward secure encryption,  $C_4$  represents the interval  $[1, 4]$ , and the private key for a user  $\omega$  can be used to derive the node secret keys for all the nodes in the interval  $[\omega, n]$  (b). Key distribution mode of interval encryption: we let  $\omega = 5$  here. The respective private key  $D_5$  contains left private key  $D_{5,(L)}$  and right private key  $D_{5,(R)}$ .  $D_{5,(L)} = \{SK_{5,(L)}, SK_{5|1,(LS),(L)}\}$  which are derived from left master key  $\alpha - \alpha_5$ . Similarly, we have  $D_{5,(R)} = \{SK_{5,(R)}, SK_{5|2,(RS),(R)}, SK_{5|3,(RS),(R)}\}$  derived from right master key  $\alpha_5$ . Let the left bound  $\eta$  of an interval be 2 here, then  $SK_{2,(L)}$  can be derived from  $SK_{2|1,(L)}$  which is equal to  $SK_{5|1,(LS),(L)}$  belonging to  $D_{5,(L)}$ .

secret keys of the leaf nodes in the interval  $[\omega, n]$  are computable from its own private key. Therefore, this private key can be used to open all the message encrypted under these nodes. Besides, we also use a left direction arrow from a leaf node  $\omega$  to denote an opposite decryption ability, namely the respective private key can be used to open all the messages encrypted under the leaf nodes in the interval  $[1, \omega]$ . It is feasible by simply assigning a user with the node secret keys for node  $\omega$  as well as node secret keys for the left siblings of all the nodes on the path from the root to  $\omega$ . Generally, the FSE construction is treated as a special interval encryption scheme in which the encryptor can set the interval form as  $[1, j]$ . The upper bound  $j$  depends on which leaf node the ciphertext corresponds to. Now, our goal is to realize an interval encryption scheme covering multiple intervals, each of which has two freely chosen bounds determined by the encryptor.

### 3.5.4 A Generic Transformation from BTE to Interval Encryption

**Trivial Constructions:** A trivial interval encryption scheme can be given directly from attribute-based encryption [129] if one treats  $\log n$  bits to represent a number from 1 to  $n$  as attributes and builds an access tree allowing specific intervals. However, even the most efficient trivial methodology would inevitably result in an interval encryption construction with a ciphertext size of  $\mathcal{O}(k \log n)$ , where  $k$  is the number of intervals. As we have mentioned, our goal is to realize a BE system in which the ciphertext size is determined by the number of intervals  $k$ . *If all the messages are only encrypted under the bounds of each interval like in the FSE scheme*, then this goal is reachable. However, it is still a challenge for us to make sure that only those with an index within two bounds of each interval can open the message.

**A Generic Transformation from BTE to Interval Encryption:** There are some difficulties in the transformation from BTE to interval encryption. The first difficulty is how to differentiate the decryption ability of an index in and outside of an interval. Taking the interval  $[3, 6]$  shown in Figure 3.12.(a), for instance, we could easily find the required difference if we project two opposite direction arrows from each index in the axis, where the connotation of the arrows can be found in our exposition of the last paragraph in Section 3.5.3. The key observation of our transformation is that: *the two opposite direction arrows starting from index 5 can cross both bounds 3 and 6, respectively, and therefore decrypt the corresponding partial ciphertext in two different manners* (We will show how to differentiate the partial decryption from two different directions, and how this will eventually lead to successful generation of the corresponding message encryption key in the sequel). *However, only one unique direction arrow from index 2 or 7 can cross the two bounds, i.e., only the right direction arrow from index 2 can cross 3 and 6 while only the left direction arrow from index 7 can cross 3 and 6. This implies that those outside of an interval can only decrypt the partial ciphertext in a unique manner.* The



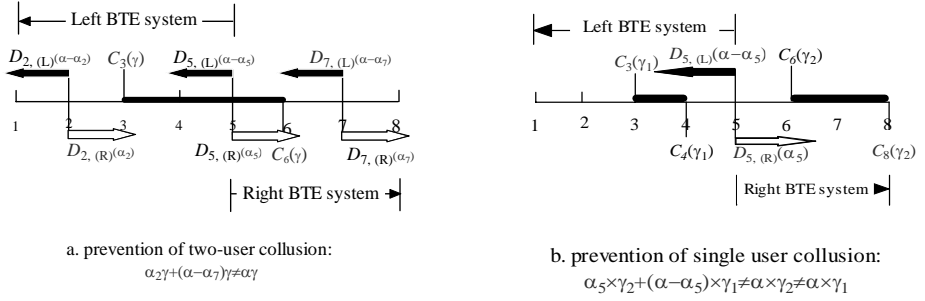


Figure 3.12: Collusion and its prevention

private key for right direction arrow is called right private key in the concrete construction and the one for left direction arrow is left private key.

The master key of the underlying BTE or HIBE scheme only contains *one group element*. The message encryption key for each interval corresponds to  $\alpha \cdot \gamma$  in the exponent of a pairing, where  $\alpha$  is the system master key and  $\gamma$  is randomly chosen by the encryptor. For each user  $\omega$ , we choose a random number  $\alpha_\omega$  and split the master key  $\alpha$  into two parts: one is the right master key  $\alpha_\omega$ , which serves as the root master key for the right BTE system, from which the right private key  $D_{\omega,(R)}$  of  $\omega$  is derived; the other one is the left master key  $\alpha - \alpha_\omega$ , i.e., the root master key for the left BTE system, from which the left private key  $D_{\omega,(L)}$  of  $\omega$  is derived. Note that the two private keys for  $\omega$  can be distributed similarly to the FSE scheme as shown in Figure 3.11. (b). Consequently, a partial decryption using the user's right private key contains  $\alpha_\omega \cdot \gamma$  in the exponent of a pairing while a partial decryption using the left private key will have  $(\alpha - \alpha_\omega) \cdot \gamma$  in its exponent. Then, the message encryption key containing  $\alpha \cdot \gamma$  in its exponent will be recovered since  $\alpha = \alpha_\omega + \alpha - \alpha_\omega$  holds.

In this way, we can actually prevent a possible collusion attack called two-user collusion. For example (shown in Figure 3.12. (a)), a user  $\omega = 7$  with a left private key  $D_{7,(L)}$  (which could decrypt the partial ciphertext  $C_3$ ) and a user  $\omega = 2$  with right private key  $D_{2,(R)}$  (which can decrypt the partial ciphertext  $C_6$ ) might collude to open the

message for interval  $[3, 6]$  (since they can also complete the partial decryption in two different manners) although neither of them is in this particular interval. In our system, the partial decryption from  $D_{7,(L)}$  will contain  $(\alpha - \alpha_7) \cdot \gamma$  while the partial decryption from  $D_{2,(R)}$  contains  $\alpha_2 \cdot \gamma$  in its exponents, and hence the collusion will fail since there's no way for them to obtain  $\alpha \cdot \gamma$  in the final step.

Besides, the encryptor is required to use a unique random  $\gamma_j$  while generating the ciphertext for each interval  $NI_j$ . This aims to prevent another attack called a single-user collusion. This attack only occurs in the scenario with multiple intervals (where  $k \geq 1$ ). For instance (shown in Figure 3.12. (b)), in an interval encryption system with two intervals  $[3, 4] \cup [6, 8]$ , the partial decryption on  $C_3$  from the left private key  $D_{5,(L)}$  contains  $\alpha - \alpha_5$  and the partial decryption on  $C_8$  from the right private key  $D_{5,(R)}$  contains the other half randomness  $\alpha_5$ , the message encryption key corresponding to  $\alpha \cdot \gamma$  might be recovered if these two intervals use the same randomness. However, a unique randomness for each interval can guarantee that only a user within a certain interval can successfully open the message. For example (as shown in Figure 3.12.(b)), a random  $\gamma_1$  is used in the ciphertext for interval  $[C_3, C_4]$  and  $\gamma_2$  is used in the encryption for interval  $[C_6, C_8]$ . The message encryption keys of these two intervals correspond to  $\alpha \cdot \gamma_1$  and  $\alpha \cdot \gamma_2$ , respectively. A single-user collusion fails since the randomized partial decryption  $(\alpha - \alpha_5)\gamma_1$  and  $\alpha_5\gamma_2$  will not generate a meaningful encryption key in the final step (see Figure 3.12. b.).

The proposed methodology might not be generic, since it somehow relies on the property of bilinear mapping. Therefore, we only illustrate our methodology using concrete examples rather than providing a formal description of a generic interval encryption system in the following subsections.

### 3.5.5 Basic Construction: A Concrete Instantiation Based on HIBE

In this subsection, we will describe how the proposed methodology can be applied to the HIBE (viewed as a binary tree encryption scheme here) construction [35] to propose an interval encryption scheme. Note that there is an additional algorithm **DeckeyDer** ( $D_\omega = \{D_{\omega,(L)}, D_{\omega,(R)}\}, \zeta, \eta$ ) compared with the original definition of interval encryption in Section 3.5.2. This algorithm is a preliminary step for the decryption algorithm, and we treat it as an independent algorithm for clarity. Besides, there is an additional slight technical modification to the underlying BTE construction in the sense that we basically have two concrete instantiations of a hash function to guarantee that we could cover both the two bounds of each interval in the security proof.

Let  $\mathbb{G}_1$  be a bilinear group of prime order  $p$ , and let  $g$  be a generator of  $\mathbb{G}_1$ . In addition, let  $\hat{e} : \mathbb{G}_1 \times \mathbb{G}_1 \rightarrow \mathbb{G}_2$  denote the bilinear map. A security parameter,  $\kappa$ , will determine the size of the groups. Assume the system accommodates  $n = 2^d$  users, where  $d$  is an integer.

**Setup**( $n$ ): Select a random  $\alpha \in \mathbb{Z}_p$  and set  $g_1 = g^\alpha$ . Choose random elements  $g_2, g_{3,(L)}, g_{3,(R)}, h_{1,(L)}, \dots, h_{d,(L)}, h_{1,(R)}, \dots, h_{d,(R)}$  from  $\mathbb{G}_1$ .

The public key is  $PK = (g, g_1, g_2, g_{3,(L)}, g_{3,(R)}, h_{1,(L)}, \dots, h_{d,(L)}, h_{1,(R)}, \dots, h_{d,(R)})$ . For a binary string  $v = [v_1 v_2 \dots v_j]$  where  $j \in [1, d]$ , define two publicly computable functions as:  $F_{(L)}(v) = g_{3,(L)} \cdot \prod_{i=1}^j h_{i,(L)}^{v_i}$  and  $F_{(R)}(v) = g_{3,(R)} \cdot \prod_{i=1}^j h_{i,(R)}^{v_i}$ . The system master key is  $SK_\varepsilon = g_2^\alpha$ .

**PvkGen**( $\omega, SK_\varepsilon$ ): For receiver  $\omega = [\omega_1 \omega_2 \dots \omega_d]$  which is associated with the  $\omega$ -th leaf node (starting from left to right), the algorithm first chooses a random number  $\alpha_\omega$ . The right master key for  $\omega$  is  $SK_{\varepsilon,(R)} = g_2^{\alpha_\omega}$ , and the left master key is  $SK_{\varepsilon,(L)} = g_2^{\alpha - \alpha_\omega}$ . The algorithm first generates two node secret keys  $SK_{\omega,(R)} = [g_2^{\alpha_\omega} (F_{(R)}(\omega))^{r_\omega}, g^{r_\omega}]$  and  $SK_{\omega,(L)} = [g_2^{\alpha - \alpha_\omega} (F_{(L)}(\omega))^{r_\omega}, g^{r_\omega}]$  for leaf node  $\omega$  where  $r_\omega$  is a random number from  $\mathbb{Z}_p$ . For all the nodes  $\omega|_j, j = 1, \dots, d$  on the path from the root to the leaf node  $\omega$ , if

it has a right sibling  $\omega|_{j,(RS)} = [\omega_1 \omega_2 \cdots \omega_{j-1} 1]$ , the algorithm uses the right master key to generate the respective node secret key as  $SK_{\omega|_{j,(RS),(R)}} = [g_2^{\alpha\omega} (F_{(R)}(\omega|_{j,(RS)}))^{r_j}, g^{r_j}, h_{j+1,(R)}^{r_j}, \dots, h_{d,(R)}^{r_j}]$  where  $r_j$  is also a random number; otherwise the algorithm uses the left master key to generate node secret key for its left sibling  $\omega|_{j,(LS)} = [\omega_1 \omega_2 \cdots \omega_{j-1} 0]$  as  $SK_{\omega|_{j,(LS),(L)}} = [g_2^{\alpha-\alpha\omega} (F_{(L)}(\omega|_{j,(LS)}))^{r_j}, g^{r_j}, h_{j+1,(L)}^{r_j}, \dots, h_{d,(L)}^{r_j}]$ .

Output private key  $D_\omega = \{D_{\omega,(R)}, D_{\omega,(L)}\}$ , where

$$D_{\omega,(R)} = \{SK_{\omega,(R)}, SK_{\omega|_{j,(RS),(R)}}\}_{j \in [1,d]}, D_{\omega,(L)} = \{SK_{\omega,(L)}, SK_{\omega|_{j,(LS),(L)}}\}_{j \in [1,d]}.$$

**Encrypt**( $S, PK$ ): The encryptor first chooses a  $k$ -wise natural interval set  $S = \bigcup_{j=1}^k NI_j$ , where  $NI_j = [l_j, r_j]$ . For each interval, pick  $\gamma_j$  uniformly from  $\mathbb{Z}_p$  at random. Let the binary name of the corresponding leaf nodes for the two bounds be  $r_j = [r_{j1} \cdots r_{jd}]$  and  $l_j = [l_{j1} \cdots l_{jd}]$ .

Output the respective ciphertext  $C_{l_j} = \{g^{\gamma_j}, (F_{(L)}(l_j))^{\gamma_j}\}$  and  $C_{r_j} = \{g^{\gamma_j}, (F_{(R)}(r_j))^{\gamma_j}\}$ . Set the message encryption key for each interval  $NI_j$  as  $K_j = \hat{e}(g_1, g_2)^{\gamma_j} \in \mathbb{G}_2$ . The collection of these partial ciphertexts constitute the header  $\text{Hdr} = \{C_{l_j}, C_{r_j}\}_{j=1}^k$ .

**DeckeyDer** ( $D_\omega = \{D_{\omega,(L)}, D_{\omega,(R)}\}, \zeta, \eta$ ): This algorithm derives the node secret key  $SK_{\eta,(L)}$  for the lower bound  $\eta$ , and  $SK_{\zeta,(R)}$  for the upper bound  $\zeta$ .

1. Let a natural number  $\eta \leq \omega$  denote the  $\eta$ -th leaf node, and thus  $\eta$  is on the left of  $\omega$  in the binary tree. Assume the binary representation of  $\eta$  is  $\eta = \eta_1 \cdots \eta_d$ . There must exist a node secret key  $SK_{\eta|_j,(L)}, j \in [1, d]$  which belongs to  $D_{\omega,(L)}$  (as shown in Figure 3.11. b.). Run the derivation algorithm of the underlying BTE scheme iteratively, which means the following steps need to be executed iteratively for  $i = j$  to  $i = d - 1$ :

(a) Let  $\eta|_i = \eta_1 \cdots \eta_i$ . Parse  $SK_{\eta|_i,(L)}$  as  $(g_2^{\alpha-\alpha\omega} (F_{(L)}(\eta|_i))^{r_i}, g^{r_i}, h_{i+1,(L)}^{r_i}, \dots, h_{d,(L)}^{r_i}) = (a_0, a_1, b_{i+1}, \dots, b_d)$ .

(b) Choose random  $t \in \mathbb{Z}_p$ , and output  $SK_{\eta|_{i+1},(L)} = (a_0 \cdot b_{i+1}^{\eta_{i+1}} \cdot (F_{(L)}(\eta|_{i+1}))^t, a_1 \cdot g^t, b_{i+2} \cdot h_{i+2,(L)}^t, \dots, b_d \cdot h_{d,(L)}^t)$  and set  $i = i + 1$ .

Finally, it will output a node secret key  $SK_{\eta,(L)} = [g_2^{\alpha-\alpha\omega} (F_{(L)}(\eta))^{r'}, g^{r'}]$  for the lower bound  $\eta$ .

2. Let a natural number  $\zeta \geq \omega$  denote the  $\zeta$ th leaf node. Assume the binary representation of  $\zeta$  is  $\zeta = \zeta_1 \cdots \zeta_d$ . Therefore, there must exist a node secret key  $SK_{\zeta|(j),(R)}$  which belongs to  $D_{\omega,(R)}$ . Run the derivation algorithm of the underlying BTE scheme iteratively, which means steps 1(a)-1(b) need to be executed iteratively.

Output a node secret key  $SK_{\zeta,(R)} = [g_2^{\alpha\omega} (F_{(R)}(\zeta))^{r''}, g^{r''}]$  for the upper bound  $\zeta$ .

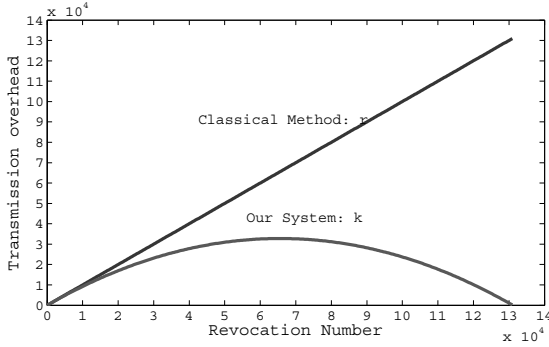
**Decrypt** ( $S, \omega, D_\omega, \text{Hdr}, PK$ ): If  $\omega \in NI_j = [l_j, r_j], 1 \leq j \leq k$  which implies that  $l_j \leq \omega \leq r_j$ , then it runs **DeckeyDer** ( $D_\omega, r_j, l_j$ ) to generate decryption key  $SK_{r_j,(R)}$  and  $SK_{l_j,(L)}$ . It obtains the corresponding secret key  $SK_{r_j,(R)} = [g_2^{\alpha\omega} (F_{(R)}(r_j))^{r''}, g^{r''}]$  and the partial ciphertext for the upper bound  $C_{r_j} = \{g^{\gamma_j}, (F_{(R)}(r_j))^{\gamma_j}\}$ . Compute  $\frac{\hat{e}[g^{\gamma_j}, g_2^{\alpha\omega} (F_{(R)}(r_j))^{r''}]}{\hat{e}[g^{r''}, (F_{(R)}(r_j))^{\gamma_j}]} = \hat{e}(g, g_2)^{\gamma_j \alpha \omega}$ . It also obtains the corresponding secret key  $SK_{l_j,(L)} = [g_2^{\alpha-\alpha\omega} (F_{(L)}(l_j))^{r'}, g^{r'}]$  and the partial ciphertext for the lower bound  $C_{l_j} = \{g^{\gamma_j}, (F_{(L)}(l_j))^{\gamma_j}\}$ . Compute  $\frac{\hat{e}[g^{\gamma_j}, g_2^{\alpha-\alpha\omega} (F_{(L)}(l_j))^{r'}]}{\hat{e}[g^{r'}, (F_{(L)}(l_j))^{\gamma_j}]} = \hat{e}(g, g_2)^{\gamma_j (\alpha-\alpha\omega)}$ . Finally, it computes  $\hat{e}(g, g_2)^{\gamma_j \alpha \omega} \cdot \hat{e}(g, g_2)^{\gamma_j (\alpha-\alpha\omega)} = \hat{e}(g^\alpha, g_2)^{\gamma_j} = \hat{e}(g_1, g_2)^{\gamma_j}$ .

### 3.5.6 Discussion on Efficiency and Security

In this construction, the public key size is  $\mathcal{O}(\log n)$ , and the private key only contains  $\mathcal{O}(\log n)$  BTE node secret keys. Note that the private key in the DF transformation [101] of the NNL method or the HS construction contains  $\mathcal{O}(\log^2 n)$  or  $\mathcal{O}(\log^{1+\epsilon} n)$  node secret keys, respectively. It is important to point out that a widely used tool, updatable public storage in the FSE scheme [35], can also be adopted in our proposed interval encryption system to limit the cost of private storage to  $\mathcal{O}(\log n)$ . The above efficiency parameters can be further improved if the random oracle is adopted, i.e., the public key size can be reduced to  $\mathcal{O}(1)$  in this case.

The decryption cost is dominated by the derivation of the two node secret keys. The derivation cost can be reduced to  $\mathcal{O}(\log n)$  by doing the following computation: in order to deduce the node secret key  $SK_{\eta, (L)} = [g_2^{\alpha - \alpha\omega} \cdot (F_{(L)}(\eta))^{r'}, g^{r'}] = (a'_0, a'_1)$  from  $SK_{\eta|_i, (L)} = (g_2^{\alpha - \alpha\omega} (F_{(L)}(\eta|_i))^{r_i}, g^{r_i}, h_{i+1, (L)}^{r_i}, \dots, h_{d, (L)}^{r_i}) = (a_0, a_1, b_{i+1}, \dots, b_d)$ , we can compute  $a'_0 = a_0 \cdot \prod_{k=i+1}^d b_k^{\eta_k} \cdot (F_{(L)}(\eta))^t$ ,  $a'_1 = a_1 \cdot g^t$  where we set  $r' = r_i + t$ . We can deduce the node secret key  $SK_{\zeta, (R)}$  from  $SK_{\zeta|_i, (R)}$  in a similar way. The overall decryption time is then reduced to  $\mathcal{O}(\log n)$  since the rest of the decryption procedure only requires a constant number of group operations.

*Why is  $\mathcal{O}(k)$  better:* A system with a transmission overhead proportional to  $k$  is more efficient than the traditional systems, especially the system where communication load is linearly dependent on  $r$  such as the revocation system [240, 303]. To illustrate, we compare the performance of both systems in presence of different values for  $r$  as well as  $k$ . We assume that the total node number  $n$  is set to  $2^{17} = 131072$  and let  $r$  increase from 1 to  $n$ . For a specific  $r$ , we randomly generate 1000 revoked sets, which correspond to 1000 different interval number  $k$ , and thus obtain an average interval number  $\bar{k}$  as well as the average transmission overhead of the proposed scheme, which has been shown in Figure 3.13. From Figure 3.13, it is observed that, when the revocation set is small, the performance of the proposed scheme is very close to the traditional systems. However the difference will be scaled up along with the increase of  $r$ . If the revoked set number exceeds 50% of the total number, the communication load of the proposed scheme will decrease with increase of  $r$ . It is also observed that the proposed scheme can achieve the best performance if  $r$  is very large, which further demonstrates that the proposed scheme is suitable for cases when a small receiver set is employed. Compared with the BGW generalized construction [39] with a  $\sqrt{n}$  size transmission overload which only serves as a better choice than the trivial solution and the traditional systems when  $r > \sqrt{n}$ , we have a benefit that our system keeps the advantage over the traditional constructions when  $r$  is a small number, namely  $r \ll n$ .

Figure 3.13: Comparison between  $k$  and  $r$ 

From the above results, it is concluded that the proposed construction fits into more cases than the traditional systems which are dependent on  $r$ , and therefore constitutes a more favorable choice in practice.

The selective security of the proposed construction can be proven secure under the  $(d+1)$ -BDHE assumption, and it is stated as follows. We omit the concrete proof here.

**Theorem 3.5.2.** *If the Decisional  $(d+1)$ -BDHE assumption holds in  $\mathbb{G}_1, \mathbb{G}_2$ , the proposed interval encryption scheme is selective chosen plaintext secure.*

### 3.5.7 Extension Work

**Inclusive Extended Interval Encryption:** An inclusive extended interval encryption scheme deals with the scenario where the message is encrypted under a collection of intervals  $S = \bigcup_{j=1}^k [l_j, r_j]$ , and the private key  $D_\omega$  of a user  $\omega$  corresponds to an interval  $[l_\omega, r_\omega]$ . The decryption is successful if at least one interval  $[l_j, r_j], j \in [1, k]$  exists such that  $[l_\omega, r_\omega] \subseteq [l_j, r_j]$ . To generate a private key corresponding to an interval  $[l_\omega, r_\omega]$  (see Figure 3.14), we simply generate a left private key  $D_{l_\omega, (L)}$  corresponding to the lower bound  $l_\omega$  using the left master key  $\alpha - \alpha_\omega$  as in the basic construction. Similarly, we generate a right private key  $D_{r_\omega, (R)}$  for the upper bound  $r_\omega$  using the right master key  $\alpha_\omega$ . The rest of the above algorithms have no significant differences from

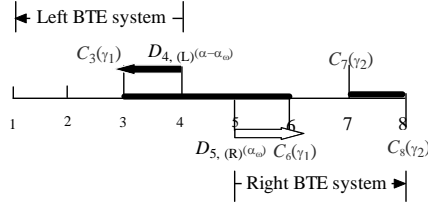


Figure 3.14: Extended Interval Encryption: The generation of a private key for an interval  $[4, 5]$

those in the basic construction. Furthermore, it is easy to observe that a man holding a private key for an interval  $[l_{\omega}, r_{\omega}]$  can delegate a private key for another interval  $[l_{\omega'}, r_{\omega'}]$  using the **DeckeyDer** algorithm as long as  $[l_{\omega}, r_{\omega}] \subseteq [l_{\omega'}, r_{\omega'}]$ . This is a property somewhat close to a recently proposed concept called inclusive identity-based encryption (IBE) [40]. We consider this extended construction of important theoretical interest because few inclusive constructions exist [120] since the proposal of inclusive IBE.

**Adaptively Secure Interval Encryption:** An adaptively secure interval encryption is constructed from Lewko and Waters’ HIBE construction [182]. The basic idea is to apply our proposed transformation method to Lewko and Waters’ HIBE scheme.

**Range Attribute-based Encryption:** In a key-policy attribute-based encryption (ABE) [129], a private key might be associated with an access policy such as “Old man **AND** tall.” A man holding this private key can open a message encrypted under an attribute set  $\{\text{“Old man,” “tall”}\}$  since this attribute set satisfies the above access policy. In practice, the attributes in an attribute set might have certain range and the attributes in an access policy might be assigned with certain concrete evaluations. In the above example, the access policy might be denoted as a formula “Age: 60 **AND** Height: 180 (cm)”. A man holding a private key associated with the above policy should be able to open a message encrypted under an attribute set  $\{\text{“Age: 50 to 100”, “Height: 175 to 250 (cm).”}\}$  The reason for the successful decryption is that both evaluations of the two



attributes fall into the range required in the attribute set and hence the access policy is satisfied. However, A man holding a private key associated with an access policy “Age: 49 **AND** Height: 180 (cm)” cannot decrypt this message since the evaluation “Age: 49” is not within the corresponding range “Age: 50 to 100” in the attribute set. A range ABE scheme is realizable from a traditional ABE scheme. However, the ciphertext will blow up with a  $\log n$  factor as shown in our trivial example of constructing interval encryption from ABE. The proposed interval encryption scheme can be easily modified to a range ABE scheme with a constant ciphertext size.

**Interval Encryption under Simpler Assumption:** The proposed method is also applicable to those BTE or HIBE schemes, in which cases their master keys only contain a single group element such as [34, 62]. We can construct interval encryption schemes based on the decisional bilinear Diffie–Hellman assumption. The concrete steps are similar to the steps of Section 3.5.5. The weakness of these constructions is that the ciphertext size will blow up with a  $\log n$  factor compared with the basic construction while the private key size remains  $\mathcal{O}(\log n)$ .

**Encryption Under a Graph:** Consider the following application: A message might be encrypted under a digital map of a certain territory on earth (a close two-dimensional graph) and only those who hold a private key for a location in the territory can open the message. This notion might actually intrigue several interesting applications. For example, a launch order of a certain weapon might be encrypted under a map of a specific region and only those who have a private key corresponding to a location within this region can launch this weapon. Apparently, we can map all the points in a two-dimensional digital map to the points in a single dimensional axis. We can simply calculate  $i = (y - 1)c + x$  where  $(x, y)$  is a point in a two-dimensional map with width  $c$  and height  $d$ . If we set  $n = c * d$ , then all the points can be mapped into an index  $i \in [1, n]$ . In other words, all the points within the territory of this digital map can be mapped into a collection of intervals. Therefore, the proposed interval encryption

provides a solution for the above scenario. The number of intervals depends on the perimeter of this graph.

**Future Work and Open Problems:** The reason why the proposed construction can reduce the transmission cost is due to the difference between the points in and outside of a certain interval. How to reduce the transmission cost based the difference between points in and outside of a graph, especially to minimize the constant factor  $k$  during the multi-dimensional scenario is left as an important open problem. It is possible to borrow some idea from computational geometry to solve this problem. To propose a BTE or HIBE construction with improved efficiency or under a weaker assumption which is compatible with our framework is very interesting since this directly implies the improvement of interval encryption.

### 3.6 Fuzzy Identity-Based Signature Schemes

At present, while we are enjoying the facilities given by the modern technology, many security and privacy issues arise. One of them is reliable user authentication. In a generic cryptographic environment, the possession of the decryption key is sufficient to authenticate the user's identity. But since the decryption key is so long and complicated, it is difficult to memorize. As a result, it is stored somewhere. It relies on other authentication (e.g., password) to retrieve this key. By simple dictionary attacks, most passwords are easy to guess. Moreover, people tend to write down complex passwords on easily accessible locations. In addition, people often use the same password in many applications. Thus, if a single password is compromised, many doors are opened.

Considering the limitation of passwords, it is not suitable to use password based authentication in systems which require high security level. An alternative approach is to use biometrics (fingerprints, iris data, face, and voice). It is inherently more reliable than passwords, since people will never forget or lost their biometrics; they are extremely difficult to copy, distribute, or share; and the person being authenticated is



Figure 3.15: Two biometric scans of the same feature are rarely identical.

required to present at the time and point of the authentication. It is also difficult to forge biometrics. Thus, biometric authentication is a potential candidate to replace password based authentication.

Despite the advantages of biometric authentication, the following issue must be addressed: biometric data is *not exactly reproducible*, as shown in Figure 3.15 that two biometric scans of the same feature are rarely identical. Thus, traditional protocols can not guarantee the correctness when the parties use a shared secret derived from biometric data.

Furthermore, a common approach to biometric authentication is to capture the biometric templates of all users during the *enrollment phase* and to store the templates in a reference database. During the *authentication phase*, new scan result is compared with the stored templates. To store biometric templates in a database will incur a number of security and privacy risks, such as:

1. Impersonation. An attacker steals templates from a database and constructs artificial biometrics that could pass authentication.
2. Irrevokability. Once compromised, biometrics cannot be updated, reissued or destroyed.
3. Exposure of sensitive personal information.

The concept of fuzzy identity-based encryption (FIBE) was introduced by Sahai and Waters and further developed in a line of works, e.g., [16, 227]. In a nutshell a FIBE

allows a user with the private key for identity  $\omega$  to decrypt a ciphertext encrypted for identity  $\omega'$  if and only if  $\omega$  and  $\omega'$  are within a certain distance judged by some metric.

In this section, we introduce a novel cryptographic primitive that is the signature analogue of a FIBE, we call it *fuzzy identity-based signature* (FIBS) [300].<sup>8</sup> FIBS allows a user with identity  $\omega$  to issue a signature which can be verified with identity  $\omega'$  if  $\omega$  and  $\omega'$  are within a certain distance judged by some metric. Using FIBS, we construct a biometric authentication scheme and address the problems which lie in biometric authentication mentioned in the above paragraph. In our scheme, the server stores a FIBS signature and a public string **pub** which is shared by the users. When a user  $A$  is going to authenticate himself to user  $B$ , he will send his public string **pub** to the server and allow the server to retrieve the signature and reconstruct the shared secret. Our scheme meets the “robust to noise,” “anti-impersonation,” and “privacy protection” security requirements.

Another interesting application is attribute-based signature. In this application, a user can issue a signature on behalf of the group that has a certain set of attributes. For example, an IT company might want a C++ senior programmer whose age is above 50 to sign the technical report. In this scenario, it will sign to the identity  $\{\text{“C++”}, \text{“senior programmer”}, \text{“above 50”}\}$ . Any user who has an identity that contains all of these attributes can issue the signature.

Much work has focused on developing secure biometric authentication [92, 164, 165, 214, 278]. Most recently, Dodis, Reyzin, and Smith [102] showed an approach different from ours to use biometric data to derive secure cryptographic keys which can be used for the purposes of authentication. Roughly speaking, they introduce two primitives: a *secure sketch* which allows recovery of a shared secret given a close approximation thereof, and a *fuzzy extractor* which extracts a uniformly string  $s$  from this shared secret in an error-tolerant manner. Both of the primitives work by constructing a “public”

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<sup>8</sup>Since FIBE appeared in 2005, we are the first to propose the cryptographic primitive FIBS in 2007.

string **pub** which is stored by the server and transmitted to the user, **pub** encodes the redundancy information needed for error-tolerant reconstruction. But the work of Dodis, Reyzin, and Smith does not address the issue of malicious modification of **pub**. Indeed, the adversary who maliciously alters the public string sent to a user may be able to learn the user's biometric data. Boyen et al. [46] improved this result in a way that resists an active adversary and provides mutual authentication and authenticated key exchange.

### 3.6.1 Definitions

**Notations:** From here on we use  $\mathbb{Z}_q$  to denote the group  $\{0, \dots, q-1\}$  under addition modulo  $q$ . For a group  $\mathbb{G}$  of prime order we use  $\mathbb{G}^*$  to denote the set  $\mathbb{G}^* = \mathbb{G} - \{O\}$  where  $O$  is the identity element in the group  $\mathbb{G}$ . We use  $\mathbb{Z}_q^*$  to denote the set of positive integers.

**Bilinear Pairings and Assumptions:** Let us consider two multiplicative group  $\mathbb{G}$  and  $\mathbb{G}_T$  of the same prime order  $p$ . A bilinear pairing is a map  $e : \mathbb{G} \times \mathbb{G} \rightarrow \mathbb{G}_T$  with the following properties:

1. Bilinear:  $e(u^a, v^b) = e(u, v)^{ab}$ , where  $u, v \in \mathbb{G}$ , and  $a, b \in \mathbb{Z}_p^*$ .
2. Nondegeneracy: there exist  $u \in \mathbb{G}$  and  $v \in \mathbb{G}$  such that  $e(u, v) \neq 1$ .
3. Computability: There is an efficient algorithm to compute  $e(u, v)$  for all  $u, v \in \mathbb{G}$ .

**Computational Diffie–Hellman (CDH) Assumption:** We briefly review the computational Diffie–Hellman (CDH) Assumption. The readers can refer to previous literature [162] for more details.

The challenger chooses  $a, b \in \mathbb{Z}_p$  at random and outputs  $(g, A = g^a, B = g^b)$ . The adversary then attempts to output  $g^{ab} \in \mathbb{G}$ . An adversary,  $\mathcal{B}$ , has at least an  $\epsilon$  advantage if

$$\Pr[\mathcal{B}(g, g^a, g^b) = g^{ab}] \geq \epsilon$$

where the probability is over the randomly chosen  $a, b$  and the random bits consumed by  $\mathcal{B}$ .

**Definition 3.6.1.** The  $(t, \epsilon)$ -CDH assumption holds if no  $t$ -time adversary has at least  $\epsilon$  advantage in winning the above game.

**Threshold Secret Sharing Schemes:** A  $(n, t)$  threshold secret sharing scheme distributes a secret  $s$  among a set of  $\mathcal{P} = \{R_1, \dots, R_n\}$  of  $n$  players by a dealer. Each player  $R_i$  will privately receive  $s_i$  as a share of the secret from the dealer. Then, those subsets with at least  $t$  players can recover the secret, while other subsets containing less than  $t$  players cannot gain any information about the secret.

*Shamir's solution* uses polynomial interpolation. Let  $GF(q)$  be a finite field with  $q \geq n$  elements, and let  $s \in GF(q)$  be the secret to be shared. The dealer randomly picks a polynomial  $f(x)$  of degree  $t - 1$ , and sets the constant of  $f(x)$  as  $s$ . So  $f(x)$  is of the following form  $f(x) = s + \sum_{j=1}^{t-1} a_j x^j$ .

If we assign every player  $R_i$  with a unique field element  $\alpha_i$ , the dealer sends the secret share  $s_i = f(\alpha_i)$  to  $R_i$  through a private channel. Now if the set of players  $S \subset \mathcal{P}$  such that  $|S| \geq t$ , they could recover the secret  $s = f(0)$  by using the following formula:

$$f(x) = \sum_{R_i \in S} \Delta_{\alpha_i, S}(x) f(\alpha_i) = \sum_{R_i \in S} \Delta_{\alpha_i, S}(x) s_i,$$

where

$$\Delta_{\alpha_i, S}(x) = \prod_{R_l \in S, l \neq i} \frac{x - \alpha_l}{\alpha_i - \alpha_l}.$$

On the other hand, it can be proved that if the subset  $B \subset \mathcal{P}$  such that  $|B| < t$ , all players in  $B$  cannot get any information about the polynomial  $f(x)$  even if they collude.

### 3.6.2 Security Models

**Fuzzy Identity-Based Signature:** The generic fuzzy identity-based signature (FIBS) scheme consists of the following algorithms:

- **Setup( $1^k, d$ ):** The Setup algorithm is a probabilistic algorithm that takes a security parameter  $1^k$  and an error tolerance parameter  $d$  as input. It generates the master key  $mk$  and public parameters  $params$ . Note that  $mk$  is kept secret.
- **Extract( $msk, ID$ ):** The private key extraction algorithm is a probabilistic algorithm that takes the master key  $mk$  and an identity  $ID$  as input. It outputs a private key associated with  $ID$ , denoted by  $D_{ID}$ .
- **Sign( $params, D_{ID}, M$ ):** The signing algorithm is a probabilistic algorithm that takes the public parameters  $params$ , a private key  $D_{ID}$  associated with  $ID$ , and a message  $M$  as input. It outputs the signature  $\sigma$ .
- **Verify( $params, ID', M, \sigma$ ):** The verification algorithm is a deterministic algorithm that takes the public parameters  $params$ , an identity  $ID'$  such that  $|ID' \cap ID| \geq d$ , the message  $M$ , and the corresponding signature  $\sigma$  as input. It returns a bit  $b$ , where  $b = 1$  means that the signature is valid.

#### Security Model

**Definition 3.6.2.** (UF-FIBS-CMA). Let  $\mathcal{A}$  be an adversary assumed to be a probabilistic Turing machine taking as input a security parameter  $k$ . Consider the following game in which  $\mathcal{A}$  interacts with a challenger  $\mathcal{C}$ :

- **Setup:** The challenger  $\mathcal{C}$  runs the setup phase of the algorithm and tells the adversary  $\mathcal{A}$  the public parameters.
- **Phase 1:**  $\mathcal{A}$  issues private key queries and signature queries for any identity  $\gamma_i$  adaptively.

- **Phase 2:**  $\mathcal{A}$  declares the target identity  $\alpha$ , where  $|\alpha \cap \gamma_i| < d$  for all  $\gamma_i$  from Phase 1.
- **Phase 3:**  $\mathcal{A}$  issues private key queries for many identities  $\gamma_j$ , where  $|\gamma_j \cap \alpha| < d$  for all  $j$ .  $\mathcal{A}$  issues signature queries for any identity.
- **Phase 4:**  $\mathcal{A}$  outputs  $(\alpha, \tilde{M}, \tilde{\sigma})$ , where  $\tilde{\sigma}$  is  $\alpha$ 's valid signature on the message  $\tilde{M}$  and  $\mathcal{A}$  does not make a signature query on  $(\tilde{M}, \tilde{\sigma})$  for identity  $\alpha$ .

We define  $\mathcal{A}$ 's success probability by

$$\text{Succ}_{\text{FIBS}, \mathcal{A}}^{\text{UF-FIBS-CMA}}(k) = \Pr[\text{Verify}(\alpha, \tilde{M}, \tilde{\sigma}) = 1]$$

The fuzzy identity-based signature scheme FIBS is said to be UF-FIBS-CMA secure if  $\text{Succ}_{\text{FIBS}, \mathcal{A}}^{\text{UF-FIBS-CMA}}(k)$  is negligible in the security parameter  $k$ .

### 3.6.3 Construction

Our scheme is based on the two level hierarchical signature in [48].

In what follows, we assume that groups  $\mathbb{G}$  and  $\mathbb{G}_T$  of prime order  $p$  such that a bilinear pairing  $e : \mathbb{G} \times \mathbb{G} \rightarrow \mathbb{G}_T$  can be constructed, and  $g$  is a generator of  $\mathbb{G}$ .

Identities will be sets of  $n$  elements of  $\mathbb{Z}_p^*$ . We use the definition of Lagrange coefficient  $\Delta_{i,S}(x)$  as in section 2.4.

**Setup(n, d):** To set up the system, first, choose  $g_1 = g^y, g_2 \in \mathbb{G}$ . Next, choose  $t_1, \dots, t_{n+1}$  uniformly at random from  $\mathbb{G}$ . Let  $N$  be the set  $\{1, \dots, n+1\}$  and we define a function,  $T$ , as:

$$T(x) = g_2^{x^n} \prod_{i=1}^{n+1} t_i^{\Delta_{i,N}(x)}.$$

Next, select a random integer  $z' \in \mathbb{Z}_p$  and a random vector  $\vec{z} = (z_1, \dots, z_m) \in \mathbb{Z}_p^m$ .



The public parameters of the system and the master key are given by

$$\begin{aligned} \mathbf{PP} &= \left( g_1, g_2, t_1, \dots, t_{n+1}, v' = g^{z'}, v_1 = g^{z_1}, \dots, v_m = g^{z_m}, \right. \\ A &= \left. e(g_1, g_2) \right) \in \mathbb{G}^{n+m+4} \times \mathbb{G}_T \\ \mathbf{MK} &= y. \end{aligned}$$

**Extract(PP, MK,  $\omega$ ):** To generate the private key for the identity  $\omega$ , first choose a random  $d - 1$  degree polynomial  $q$  such that  $q(0) = y$ , and return

$$K_\omega = (\{D_i\}_{i \in \omega}, \{d_i\}_{i \in \omega}) \in \mathbb{G}^{2n},$$

where the elements are constructed as

$$D_i = g_2^{q(i)} T(i)^{r_i}, \quad d_i = g^{-r_i},$$

where  $r_i$  is a random number from  $\mathbb{Z}_p$  defined for all  $i \in \omega$ .

**Sign(PP,  $K_\omega$ ,  $M$ ):** To sign a message represented as a bit string  $M = (\mu_1 \cdots \mu_m) \in \{0, 1\}^m$  for identity  $\omega$ , using private key  $K_\omega = (\{D_i\}_{i \in \omega}, \{d_i\}_{i \in \omega}) \in \mathbb{G}^{2n}$ , select a random  $s_i \in \mathbb{Z}_p$  for each  $i$  in  $\omega$ , and output

$$\begin{aligned} S &= \left( \left\{ D_i \cdot \left( v' \prod_{j=1}^m v_j^{\mu_j} \right)^{s_i} \right\}_{i \in \omega}, \{d_i\}_{i \in \omega}, \{g^{-s_i}\}_{i \in \omega} \right) \\ &= \left( \left\{ g_2^{q(i)} \cdot T(i)^{r_i} \cdot \left( v' \prod_{j=1}^m v_j^{\mu_j} \right)^{s_i} \right\}_{i \in \omega}, \{g^{-r_i}\}_{i \in \omega}, \{g^{-s_i}\}_{i \in \omega} \right) \in \mathbb{G}^{3n}. \end{aligned}$$

**Verify(PP,  $\omega'$ ,  $\mathbf{M}$ ,  $\sigma$ ):** To verify a signature  $S = (\{S_1^{(i)}\}_{i \in \omega}, \{S_2^{(i)}\}_{i \in \omega}, \{S_3^{(i)}\}_{i \in \omega})$  with respect to an identity  $\omega'$ , where  $|\omega' \cap \omega| \geq d$ , and a message  $M = (\mu_1, \dots, \mu_m) \in$

$\{0, 1\}^m$ , choose an arbitrary  $d$ -element subset  $S$  of  $\omega \cap \omega'$  and verify that

$$\begin{aligned}
& \prod_S \left( e \left( S_1^{(i)}, g \right) \cdot e \left( S_2^{(i)}, T(i) \right) \cdot e \left( S_3^{(i)}, v' \prod_{j=1}^m v_j^{\mu_j} \right) \right)^{\Delta_{i,S}(0)} \\
&= \prod_S \left( e \left( g_2^{q(i)} \cdot T(i)^{r_i} \cdot \left( v' \prod_{j=1}^m v_j^{\mu_j} \right)^{s_i}, g \right) \cdot e \left( g^{-r_i}, T(i) \right) \right. \\
&\quad \left. \cdot e \left( g^{-s_i}, v' \prod_{j=1}^m v_j^{\mu_j} \right) \right)^{\Delta_{i,S}(0)} \\
&= \prod_S \left( e \left( g_2^{q(i)}, g \right) \cdot e \left( T(i)^{r_i}, g \right) \cdot e \left( \left( v' \prod_{j=1}^m v_j^{\mu_j} \right)^{s_i}, g \right) \cdot e \left( g^{-r_i}, T(i) \right) \right. \\
&\quad \left. \cdot e \left( g^{-s_i}, v' \prod_{j=1}^m v_j^{\mu_j} \right) \right)^{\Delta_{i,S}(0)} \\
&= \prod_S e \left( g_2^{q(i)}, g \right)^{\Delta_{i,S}(0)} \stackrel{?}{=} A.
\end{aligned}$$

If the equality holds, output **1**; otherwise, output **0**.

### 3.6.4 Security Proofs

We show security as in Theorem 5.1, the approach is based on that of [48, 129].

**Theorem 3.6.3.** *Let  $\mathcal{A}$  be an adversary that makes at most  $l \ll p$  signature queries and produces a successful forgery against our scheme with probability  $\epsilon$  in time  $t$ . Then there exists an algorithm  $\mathcal{B}$  that solves the CDH problem in  $\mathbb{Z}_p$  with probability  $\tilde{\epsilon} \geq \epsilon/(4p^3nl)$  in time  $\tilde{t} \approx t$ .*

*Proof.* The simulator  $\mathcal{B}$  is given an instance  $(g, g^a, g^b) \in \mathbb{G}^3$  of the CDH problem, and wants to produce  $g^{ab}$ . The simulation proceeds as follows:

**Setup:**  $\mathcal{B}$  first selects a random identity  $\alpha^*$ . Next,  $\mathcal{B}$  chooses a random  $k \in \{0, \dots, m\}$ , and random numbers  $x', x_1, \dots, x_m$  in the interval  $\{0, \dots, 2l - 1\}$ . It also chooses

additional random exponents  $z', z_1, \dots, z_m \in \mathbb{Z}_p$ . It lets  $g_1 = g^a, g_2 = g^b$ . It then chooses a random  $n$  degree polynomial  $f(x)$  and an  $n$  degree polynomial  $u(x)$  such that  $\forall x \ u(x) = -x^n$  if  $x \in \alpha$ .  $\mathcal{B}$  sets  $t_i = g_2^{u(i)} g^{f(i)}$  for  $i$  from 1 to  $n+1$ . Since  $t_i$  is chosen independently at random, we have  $T(i) = g_2^{i^n} \prod_{j=1}^{n+1} (g_2^{u(j)} g^{f(j)})^{\Delta_{j,N}(i)} = g_2^{i^n + u(i)} g^{f(i)}$ . The simulator gives the public parameters,

$$\mathbf{PP} = (g, g_1, g_2, t_1, \dots, t_{n+1}, v' = g_2^{x' - 2kl} g^{z'}, \{v_j = g_2^{x_j} g^{z_j}\}_{j=1, \dots, m}, A = e(g_1, g_2)).$$

The corresponding master key,  $\mathbf{MK} = a$ , is unknown to  $\mathcal{B}$ .

To answer a private key query on identity  $\gamma$  that  $|\gamma \cap \alpha^*| < d$ , the simulator  $\mathcal{B}$  proceeds as follows. We first define three sets  $\Gamma, \Gamma', S$  in the following manner:

$$\Gamma = \gamma \cap \alpha, \Gamma' \text{ be any set such that } \Gamma \subseteq \Gamma' \subseteq \gamma \text{ and, } |\Gamma'| = d-1, \text{ and } S = \Gamma' \cup \{0\}.$$

Then we define the private key  $K_\gamma$  for  $i \in \Gamma'$  as  $(\{D_i\}_{i \in \Gamma'} = \{g_2^{\lambda_i} T(i)^{r_i}\}_{i \in \Gamma'}, \{d_i\}_{i \in \Gamma'} = \{g^{r_i}\}_{i \in \Gamma'})$ , where  $\lambda_i, r_i$  are chosen randomly in  $\mathbb{Z}_p$ . We define  $d-1$  degree polynomial  $q(x)$  as  $q(i) = \lambda_i, q(0) = a$ .

Next we compute the private key  $K_\gamma$  for  $i \in \gamma - \Gamma'$  as follows:

$$\begin{aligned} D_i &= \left( \prod_{j \in \Gamma'} g_2^{\lambda_j \Delta_{j,S}(i)} \right) \cdot \left( g_1^{\frac{-f(i)}{i^n + u(i)}} \left( g_2^{i^n + u(i)} g^{f(i)} \right)^{r_i'} \right)^{\Delta_{0,S}(i)} \\ d_i &= \left( g_1^{\frac{-1}{i^n + u(i)}} g^{r_i'} \right)^{\Delta_{0,S}(i)}. \end{aligned}$$

Since  $i \notin \alpha$ ,  $i^n + u(i)$  can not be zero. We claim that such construction is a valid response to this private key query. To see this, let  $r_i = (r_i' - \frac{a}{i^n + u(i)}) \Delta_{0,S}(i)$ . Then we

have that

$$\begin{aligned}
D_i &= \left( \prod_{j \in \Gamma'} g_2^{\lambda_j \Delta_{j,S(i)}} \right) \cdot \left( g_1^{\frac{-f(i)}{i^n+u(i)}} (g_2^{i^n+u(i)} g^{f(i)})^{r_i'} \right)^{\Delta_{0,S(i)}} \\
&= \left( \prod_{j \in \Gamma'} g_2^{\lambda_j \Delta_{j,S(i)}} \right) \cdot \left( g^{\frac{-af(i)}{i^n+u(i)}} (g_2^{i^n+u(i)} g^{f(i)})^{r_i'} \right)^{\Delta_{0,S(i)}} \\
&= \left( \prod_{j \in \Gamma'} g_2^{\lambda_j \Delta_{j,S(i)}} \right) \cdot \left( g_2^a (g_2^{i^n+u(i)} g^{f(i)})^{\frac{-a}{i^n+u(i)}} \cdot (g_2^{i^n+u(i)} g^{f(i)})^{r_i'} \right)^{\Delta_{0,S(i)}} \\
&= \left( \prod_{j \in \Gamma'} g_2^{\lambda_j \Delta_{j,S(i)}} \right) \cdot \left( g_2^a (g_2^{i^n+u(i)} g^{f(i)})^{r_i' - \frac{a}{i^n+u(i)}} \right)^{\Delta_{0,S(i)}} \\
&= \left( \prod_{j \in \Gamma'} g_2^{\lambda_j \Delta_{j,S(i)}} \right) g_2^{a \Delta_{0,S(i)}} (T(i))^{r_i} \\
&= g_2^{q(i)} T_i^{r_i} \\
d_i &= \left( g_1^{\frac{-1}{i^n+u(i)}} g^{r_i'} \right)^{\Delta_{0,S(i)}} = \left( g^{r_i' - \frac{a}{i^n+u(i)}} \right)^{\Delta_{0,S(i)}}.
\end{aligned}$$

It shows that  $D_i, d_i$  have the correct distribution. To answer the signature query on identity  $\gamma$  that  $|\gamma \cap \alpha^*| < d$ ,  $\mathcal{B}$  uses  $K_\gamma$  to create a signature on  $M$  exactly as in the actual scheme, and outputs the result.

To answer the signature query on identity  $\alpha^*$  for some  $M = (\mu_1 \cdots \mu_m)$ , we define  $F = -2kl + x' + \sum_{j=1}^m x_j \mu_j$  and  $J = z' + \sum_{j=1}^m z_j \mu_j$ . If  $F \equiv 0 \pmod{p}$ , the simulator aborts. Otherwise,  $\mathcal{B}$  selects a random set  $\Lambda$  such that  $\Lambda \subset \alpha^*$  and  $|\Lambda| = d - 1$  and define  $g^{q'(i)} = g^{\lambda'_i}$  for  $i \in \Lambda$  where  $\lambda'_i$  is chosen randomly in  $\mathbb{Z}_p$ . Then it computes  $g^{q'(i)} = (\prod_{j=1}^{d-1} g^{\lambda'_j \Delta_{j,\alpha^*}(i)}) g^{a \Delta_{0,\alpha^*}(i)}$  for  $i \in \alpha^* - \Lambda$ .  $\mathcal{B}$  picks random  $r_i, s_i$  for  $i \in \alpha^*$  and computes,

$$S_1^{(i)} = (g^{q'(i)})^{-J/F} g^{f(i)r_i} (g^J g_2^F)^{s_i}$$

$$S_2^{(i)} = g^{-r_i}$$

$$S_3^{(i)} = (g^{q'(i)})^{1/F} g^{-s_i}.$$

For  $\tilde{s}_i = s_i - q'(i)/F$ , we have

$$\begin{aligned}
 S_1^{(i)} &= (g^{q'(i)})^{-J/F} g^{f(i)r_i} (g^J g_2^F)^{s_i} \\
 &= (g^{q'(i)})^{-J/F} g^{f(i)r_i} g^{q'(i)J/F} g_2^{q'(i)} (g^J g_2^F)^{s_i - q'(i)/F} \\
 &= g_2^{q'(i)} g^{f(i)r_i} (g^J g_2^F)^{\tilde{s}_i} = g_2^{q'(i)} T(i)^{r(i)} (v' \prod_{j=1}^m v_j^{\mu_j})^{\tilde{s}_i} \\
 S_3^{(i)} &= (g^{q'(i)})^{1/F} g^{-s_i} = (g^{q'(i)})^{1/F} g^{-q'(i)/F} g^{-\tilde{s}_i} = g^{-\tilde{s}_i}.
 \end{aligned}$$

It shows that  $S_1^{(i)}$ ,  $S_2^{(i)}$ , and  $S_3^{(i)}$  have the correct distribution. Eventually,  $\mathcal{A}$  outputs a valid forgery  $S^* = (\{S_1^{(i)*}\}_{i \in \alpha}, \{S_2^{(i)*}\}_{i \in \alpha}, \{S_3^{(i)*}\}_{i \in \alpha})$  on  $M^*$  where  $M^* = (\mu_1^* \cdot \dots \mu_m^*) \in \{0, 1\}^m$  for identity  $\alpha$ . Let  $F^* = -2kl + x' + \sum_{j=1}^m x_j \mu_j^*$  and  $J^* = z' + \sum_{j=1}^m z_j \mu_j^*$ . If  $\alpha \neq \alpha^*$  or  $F^* \not\equiv 0 \pmod{p}$ ,  $\mathcal{B}$  aborts. Otherwise, the forgery must be of the following form, for some  $r_i^*, s_i^* \in \mathbb{Z}_p$ ,

$$\begin{aligned}
 S_1^{(i)} &= g_2^{q^*(i)} T(i)^{r_i^*} \left( v' \prod_{j=1}^m v_j^{\mu_j} \right)^{s_i^*} = g_2^{q^*(i)} g^{f(i)r_i^*} g^{J^* s_i^*} \\
 S_2^{(i)} &= g^{-r_i^*} \\
 S_3^{(i)} &= g^{-s_i^*}.
 \end{aligned}$$

We select a random set  $\Lambda'$  such that  $\Lambda' \subset \alpha$  and  $|\Lambda'| = d$ , and compute as follows:

$$\begin{aligned}
 S_1^* &= \prod_{i \in \Lambda'} \left( S_1^{(i)} \right)^{\Delta_{i, \alpha}(i)} \\
 &= \prod_{i \in \Lambda'} \left( g_2^{\Delta_{i, \alpha}(i) q^*(i)} T(i)^{\Delta_{i, \alpha}(i) r_i^*} \cdot \left( v' \prod_{j=1}^m v_j^{\mu_j} \right)^{\Delta_{i, \alpha}(i) s_i^*} \right) \\
 &= \prod_{i \in \Lambda'} \left( g_2^{\Delta_{i, \alpha}(i) q^*(i)} g^{\Delta_{i, \alpha}(i) f(i) r_i^*} g^{\Delta_{i, \alpha}(i) J^* s_i^*} \right) \\
 &= g^{ab} \prod_{i \in \Lambda'} \left( g^{\Delta_{i, \alpha}(i) f(i) r_i^*} g^{\Delta_{i, \alpha}(i) J^* s_i^*} \right)
 \end{aligned}$$

$$\begin{aligned}
S_2^* &= \prod_{i \in \Lambda'} \left( S_2^{(i)} \right)^{\Delta_{i,\alpha}(i)f(i)} = \prod_{i \in \Lambda'} g^{-\Delta_{i,\alpha}(i)f(i)r_i^*} \\
S_3^* &= \prod_{i \in \Lambda'} \left( S_3^{(i)} \right)^{\Delta_{i,\alpha}(i)} = \prod_{i \in \Lambda'} g^{-\Delta_{i,\alpha}(i)s_i^*}.
\end{aligned}$$

$\mathcal{B}$  could solve the CDH instance by outputting  $S_1^* \cdot S_2^* \cdot (S_3^*)^{J^*} = g^{ab}$ .

$$\begin{aligned}
&\Pr[\text{the simulation not aborting}] \\
&= \Pr[\alpha = \alpha^*] \cdot \Pr[F \not\equiv 0 \pmod{p}] \cdot \Pr[F^* \equiv 0 \pmod{p}] \\
&= \frac{1}{p^n} \cdot \left( 1 - \frac{1}{2l} \right) \cdot \frac{1}{2nl} \leq \frac{1}{4p^n nl}
\end{aligned}$$

Thus,  $\tilde{\epsilon} \geq \epsilon \cdot \Pr[\text{the simulation not aborting}] \geq \epsilon \cdot \frac{1}{4p^n nl}$ . □

### 3.6.5 Applications to Biometric Authentication

An important application of FIBS deals with biometric authentication. A biometric authentication system is essentially a pattern-recognition system that recognizes a person based on a feature vector derived from a specific physiological or behavioral characteristic that the person possesses [228]. Since biometrics cannot be lost or forgotten like e.g., computer passwords, biometrics have the potential to offer higher security and more convenience for the users.

Our FIBS system provides an attractive solution to biometric authentication. We use our FIBS scheme (Setup, Extract, Sign, Verify) presented in Section 3.6.3 as the underlying algorithm. We adopt the biometric authentication model presented by Verbitskiy *et al.* It consists of two phases: An enrollment phase and an authentication / Verification phase. Figure 3.16 gives an illustration of our protocol. In Figure 3.16, the text above the arrow represents the sending message. For instance, 1. X means in step 1, Alice's biometric measurement data X is sent to Extract Box. Our protocol performs as follows:

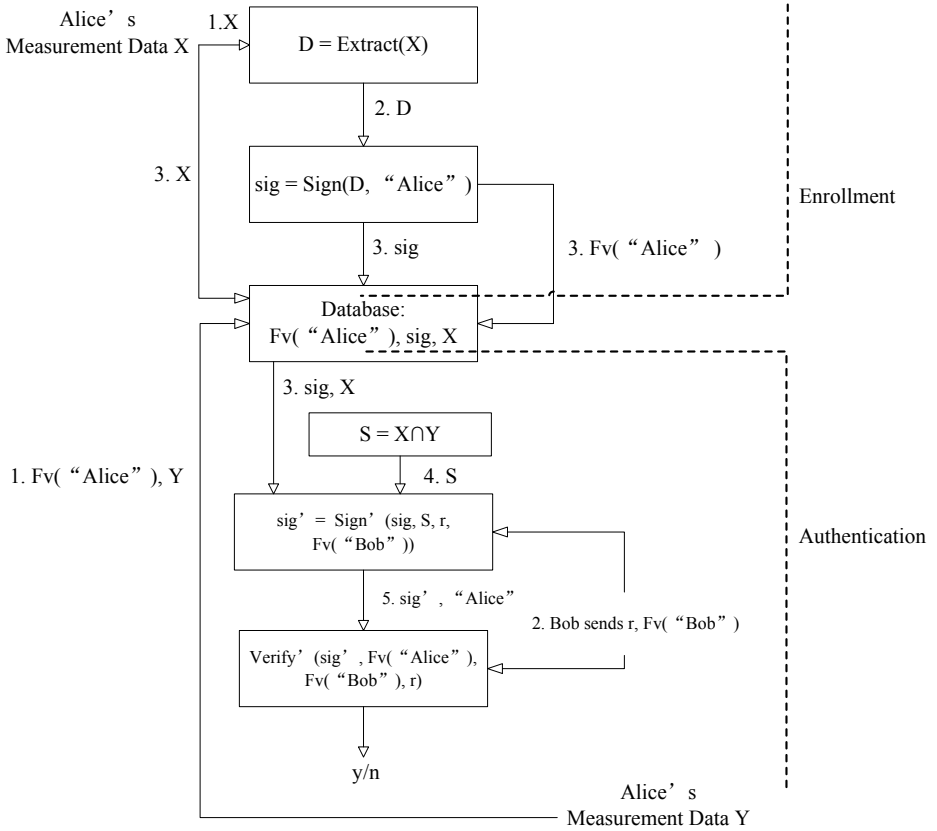


Figure 3.16: The proposed biometric authentication architecture.

**Enrollment Phase:**

1. First, Alice goes with her biometric data through an enrollment phase to a certification authority (CA). During this procedure the properties of her biometric data are measured with special equipment. We model the measurement data as a feature vector, which is a set of  $n$  elements of  $Z_p^*$ .
2. From the measurement data a private key  $D$  is derived using **Extract**.
3. Then the reference data stored in the database is obtained by applying **Sign** to  $D$  using "Alice" as the message. CA stores  $F_v(\text{"Alice"})$  together with the reference

data and the biometric measurement data. Here  $F_v(M) = v' \prod_{j=1}^m v_j^{\mu_j}$ ,  $M = (\mu_1, \dots, \mu_m) \in \{0, 1\}^m$ , and  $v', v_1, \dots, v_m$  are public parameters.

4. CA then erases the private key  $D$ , the message “Alice,” and all the intermediate data physically.

#### Authentication / Verification Phase:

1. When Alice wants to authenticate herself to Bob at a later point of time, a measurement that extracts her biometric data  $Y$  is taken. Alice sends  $F_v(\text{“Alice”})$  and  $Y$  to CA.
2. Bob sends  $F_v(\text{“Bob”})$  and a random number  $r$  to CA.
3. CA find the reference data  $\sigma = \{\sigma_1^{(i)}, \sigma_2^{(i)}, \sigma_3^{(i)}\}$  and biometric measurement data  $X$  by searching  $F_v(\text{“Alice”})$ .
4. CA computes  $S = X \cap Y$ .
5. CA computes  $\sigma^* = \{\sigma_1^{*(i)}, \sigma_2^{*(i)}, \sigma_3^{*(i)}, \sigma_4^{*(i)}\}_{i \in S} = \{(\sigma_1^{(i)} F_v(\text{“Bob”}))^{\Delta_{i,S}(0)r}, (\sigma_2^{(i)})^{\Delta_{i,S}(0)r}, (\sigma_3^{(i)})^{\Delta_{i,S}(0)r}, g^{\Delta_{i,S}(0)r}\}_{i \in S}$ , and returns  $\sigma^*$  to Alice. Here  $T(i)$  is the public function defined by our FIBS scheme and  $S$ ’s description is included in  $\sigma^*$ . Then, Alice sends  $\sigma^*$  and her identity “Alice” to Bob.
6. Taking  $\sigma^*$  and “Alice” as input parameters, Bob checks whether

$$\prod_{i \in S} (\sigma_1^{*(i)}, g) \cdot e(\sigma_2^{*(i)}, T(i)) \cdot e(\sigma_3^{*(i)}, F_v(\text{“Alice”})) e(\sigma_4^{*(i)}, F_v(\text{“Bob”})) \stackrel{?}{=} A^r,$$

$A$  is public parameters. If it holds, Alice passes the authentication. Otherwise, Alice fails the authentication.

**Security Analysis:** The security of our protocol is analyzed in terms of three important security measurements:



1. “Robust to noise”: Biometric data is *not exactly reproducible*, as two biometric scans of the same feature are rarely identical.
2. “Anti-Impersonation”: An attacker steals templates from a database and constructs artificial biometrics that pass authentication.
3. “Privacy protection”: Exposure of sensitive personal information.

**Analysis of “Robust to noise”** : With the proposed protocol, the Setup algorithm is run to set up the error tolerance parameter  $d$  according to the noise between two biometric scans. The reference data  $\sigma$  stored in the database is generated using  $d$ , such that the people with the biometric measurement  $\omega'$  will be authenticated as  $\omega$  if  $\omega$  and  $\omega'$  is within the distance  $d$ . Therefore, our protocol meets the “Robust to noise” requirements.

**Analysis of “Anti-Impersonation”**: This security requirement was recognized by several authors [33, 274, 275]. When an authentication system is used on a large scale, the reference database has to be made available to many different verifiers who cannot be trusted. Especially in the network environment, this is a serious threat. It was explicitly shown [211] that by using information stolen from a database, artificial biometrics can be constructed to impersonate people. Construction of artificial fingerprints is possible even if part of the template is available. It has been shown [141] that if minutiae templates of a fingerprint are available, it is still possible to construct artificial fingerprints that pass the authentication.

The proposed protocol satisfies “Anti-Impersonation” requirement because the underlying FIBS scheme is proved secure under UF-FIBS-CMA model. In UF-FIBS-CMA model, the adversary’s probability to produce a valid signature of any message under a new private key is negligible, even though he can get any number of signatures and private keys. In our protocol, the reference data stored in the database is a FIBS signature. Therefore, if the adversary can construct an artificial template that is different from the template stored in the database, he can provide a successful forgery of the

underlying FIBS scheme.

**Analysis of “Privacy protection”:** It is shown [228] that several privacy concerns surround the use of biometrics for authentication. Because biometric identifiers are biological in origin, collectors may glean additional personal information from scanned biometric measurements. For instance, certain malformed fingers might be statistically correlated with certain genetic disorders. With the rapid advances in human genome research, the fear of inferring further information from biological measurements becomes serious.

Because the identity data “Alice” stored in the database is  $F_v(\text{“Alice”})$  and the discrete logarithm problem is difficult in group  $\mathbb{G}$ , the probability to recover “Alice” is negligible. At the same time, the authentication request sent to CA only contains  $F_v(\text{“Alice”})$ . Therefore, the adversary cannot get any relationship between scanned biometric measurement and identity information stored in the database. Therefore, our protocol meet the “privacy protection” requirements.

**Efficiency** We now consider the efficiency of the scheme in terms of data size stored in the database and computation time in enrollment phase and authentication/verification phase.

Let  $l_{\mathbb{G}}$  and  $l_{Z_p^*}$  be the bit length of elements in group  $\mathbb{G}$  and  $Z_p^*$  respectively. And the biometric measurement data contains  $n$  elements in  $Z_p^*$ . The signature consists of three group elements in  $\mathbb{G}$  for every element in feature vector  $\omega$ . That is, the length of the signature is  $3nl_{\mathbb{G}}$ . The length of  $F_v(m)$  is  $l_{\mathbb{G}}$ . The length of biometric measurement data is  $nl_{Z_p^*}$ . Therefore, The total length of each data is  $3nl_{\mathbb{G}} + nl_{Z_p^*} + l_{\mathbb{G}}$ .

The computation time in enrollment phase is  $m + n$  multiplication in  $\mathbb{G}$  and  $2n$  exponentiation computation in  $\mathbb{G}$ .

In our rudimentary Sign algorithm, the number of pairings to sign might be as large as the number of elements of the biometric measurement. However, it is possible to

optimize this method. We now discuss methods to improve upon it.

We find out that

$$e(S_1^{(i)}, g), e(S_2^{(i)}, T(i)), e(S_3^{(i)}, v' \prod_{j=1}^m v_j^{\mu_j})$$

can be computed offline. Therefore, we compute and store

$$e(S_1^{(i)}, g) \cdot e(S_2^{(i)}, T(i)) \cdot e(S_3^{(i)}, v' \prod_{j=1}^m v_j^{\mu_j})$$

to replace the original signature stored in the database. There will be  $3n$  more pairing computations in the enrollment phase. So the computation time for enrollment phase is  $m + n$  multiplications in  $\mathbb{G}$ ,  $2n$  exponentiation computations in  $\mathbb{G}$ ,  $3n$  pairing computations. And the total length of each data will be  $nl_{\mathbb{G}_T} + nl_{Z_p^*} + l_{\mathbb{G}}$ . The computation time in the authentication/verification phase is drastically reduced to  $n$  exponentiation computations in  $\mathbb{G}$ ,  $2n$  exponentiation computations in  $\mathbb{G}_T$ ,  $2(d - 1)$  multiplications in  $\mathbb{G}_T$ .

It is shown [97], the data length of iris biometric template is of 150 dimensions. Therefore,  $n$  is 150. And  $l_{Z_p^*}$  is 1024 bits, while  $l_{\mathbb{G}_T}$  and  $l_{\mathbb{G}}$  are 512 bits in normal case. In summary, we have  $nl_{\mathbb{G}_T} + nl_{Z_p^*} + l_{\mathbb{G}}$  bit length, which is nearly 28KB for each data stored in the database.

Furthermore, we observe that the computation for enrollment phase could be run concurrently, so the total computation time for Enrollment phase is roughly  $3n$  pairing computation. We consider the benchmarks of the primitive cryptographic operations given by Jiang et al. [161], which is evaluated on Intel Core™ 2 Duo 1.83 GHz Linux machine. The computation time for pairing is 2.75ms and multiplication is 0.75ms. So the enrollment phase takes less than 1.25s, and the authentication/verification phase takes less than 0.62s. Note that a pairing operation is roughly two to three times more

expensive than exponentiation in  $\mathbb{G}$  and  $\mathbb{G}_T$ .

We have demonstrated that the proposed protocol can function efficiently both in the authentication/verification phase and the Enrollment phase.

In the recent work, ABS has been a hot topic [185,220]. In 2012, two attribute-based signature schemes with constant size signatures are proposed [139]. Their security is proven in the selective-predicate and adaptive-message setting, in the standard model, under chosen message attacks, with respect to some algorithmic assumptions related to bilinear groups.

### 3.7 Notes

In this chapter, we have given a comprehensive description of the formal definitions and security models of attribute-based encryption cryptosystems. We also introduced some of our research work publicized recently in this field, namely a bounded ciphertext-policy encryption scheme published in the conference ASIACCS 2009, a multi-authority attribute-based encryption scheme published in the conference Indocrypt 2008 and ESORICS 2011, an interval encryption scheme published in the conference ACNS 2010 and a FIBS scheme published in the journal *Computers & Electrical Engineering*, which has become the attribute-based signature in fact. Readers can refer to [192, 198, 199, 201, 301], if they want to know more ABE schemes and ABS schemes.

There are also some interesting problems in attribute-based cryptography. For instance,

1. How to design an expressive CPABE system resisting malicious key leakage?

*For the bounded ciphertext-policy attribute-based encryption, it is imperative for improving the efficiency of the proposed BCPABE scheme and designing a new kind of BCPABE excluding the assistance of dummy nodes. In a CPABE system, decryption keys are defined over*

*attributes shared by multiple users. Given a decryption key it may not be possible to trace to the original key owner. Also, as a decryption privilege could be possessed by multiple users who own the same set of attributes, malicious users would tempt to or be very willing to leak their decryption privileges without the risk of being caught. This problem severely limits the applications of CPABE. Although some attempts [142, 186, 187] have been made to address this problem, the policy in these systems can only support a single AND gate with wild card. To construct an expressive CPABE system resisting malicious key leakage is still an open problem.*

2. How to design interval encryption schemes in the multi-dimensional scenario?
3. How to design attribute-based proxy re-encryption schemes with higher efficiency and stronger security?

*Combining the proxy re-encryption technique with recently introduced attribute-based cryptosystem, attribute-based proxy re-encryption scheme is proposed. All the advantages of PRE scheme can be inherited into access control environment. The security model of ABPRE was defined for the first time and our scheme can be proved selective-structure chosen plaintext secure and master key secure in the standard model assuming that ADBDH problem and ADH problem are hard to solve. Moreover, another kind of key delegation algorithm is developed in the ABPRE scheme, providing more delegating capability to each valid user. The future work includes how to design a more delicate ABPRE scheme with higher efficiency and stronger security. Besides, it still remains an open problem to construct an ABPRE scheme that has a reduction based on a more standard and natural assumption.*

# Chapter 4

## Batch Cryptography

### 4.1 Introduction

From Chapter 1, we know the new direction of modern cryptography — *Batch Cryptography*, which can process the decryption, key agreement, and signature/verification in a batch way, instead of one by one. In this chapter, we give the formal definitions and concrete schemes as well as the related theories in mathematics. Batch cryptography focuses on the global efficiency and security. It includes batch decryption, batch key agreement and batch signature/verification. To the best of our knowledge, it is the first effort towards introducing batch cryptography.

The remainder of this chapter is organized as follows. A brief review on related work of batch signature/verification schemes as well as their security models and schemes are given in Section 4.2. Batch decryption and batch key agreement are introduced in Section 4.3. Batch RSA is considered as an important case of batch cryptography. New batch RSA schemes based on Diophantine equations are presented in Section 4.4. To implement these new schemes, we employ some related mathematical theory and techniques for solving these Diophantine equations in Section 4.5. Finally, a conclusion

of the chapter is presented in Section 4.6.

## 4.2 Aggregate Signature and Batch Verification

Aggregate signature was first introduced by Boneh et al. [38] for batch verification, which was based on the short signature [41] with efficiently computable bilinear maps in 2003. An aggregate signature scheme is a special digital signature that supports aggregation which can be informally defined as follows: given  $b$  signatures on  $b$  distinct messages from  $b$  distinct users, it is possible to aggregate all these signatures  $\sigma_1, \sigma_2, \dots, \sigma_b$  into a single *compact* signature which convinces the verifier that the  $b$  users indeed signed the  $b$  original messages in a batch way (i.e., user  $i$  signed message  $m_i$  for  $i = 1, 2, \dots, b$ ). Moreover, the aggregation can be performed incrementally. That is, signatures  $\sigma_1, \sigma_2$  can be aggregated into  $\sigma_{12}$  which can then be further aggregated with  $\sigma_3$  to obtain  $\sigma_{123}$ .

$$\sigma_1 + \sigma_2 \rightarrow \sigma_{12} \text{ and } \sigma_{12} + \sigma_3 \rightarrow \sigma_{123}.$$

Lysyanskaya et al. [205] use certified trapdoor permutations which permits only sequential aggregation, i.e., the  $b$ -th signer must aggregate its own signature into the aggregate signature formed by the first  $b - 1$  signers. Lu et al. [203] present the first aggregate signature scheme that is provably secure without random oracles. The signatures are sequentially constructed. However, unlike the scheme of Lysyanskaya et al. [205], a verifier need not know the order in which the aggregate signature was created. Additionally, these signatures are shorter than those of Lysyanskaya et al. [205] and can be verified more efficiently than those of Boneh et al. [38]. Ahn et al. [3] modify the stateful signatures of Hohenberger and Waters [146] by removing the chameleon hash and present a surprisingly efficient synchronized aggregate signature scheme which is secure under the computational Diffie–Hellman (CDH) assumption in the standard model.

In the batch verification, Naccache et al. [219] gave the first efficient batch verifier for DSA signatures in 1994; however, an interactive batch verifier presented in an early version of their paper was broken by Lim and Lee [196]. In 1995, Lai and Yen [302] proposed a new method for batch verification of DSA and RSA signatures, but the RSA batch verifier was broken five years later by Boyd and Pavlovski [45]. In 1998, Harn presented two batch verification techniques for DSA [137] and RSA [138] but both were later broken [45, 152, 153]. In the same year, Bellare et al. [21] took the first systematic look at batch verification and presented three generic methods for batching modular exponentiations, called the random subset test, the small exponents test, and the bucket test which are similar to the ideas from Lai and Yen [302]. Later, Cheon and Lee [83] introduced two new methods called the sparse exponents test and the complex exponents test, which they claim to be about twice as fast as the small exponents test. In 2000, Boyd and Pavlovski [45] proposed some attacks against different batch verification schemes most of which were based on the small exponents test and related tests, and repaired some broken schemes based on the small exponents test. In 2001, Hoshino et al. [148] pointed out that the problem discovered by Boyd and Pavlovski [45] was only critical for batch verification such as the zero-knowledge proofs. Other schemes for batch verification based on bilinear maps were proposed [304, 308, 309] but all were later broken by Cao et al. [67]. In 2006, a method was proposed for identifying invalid signatures in RSA-type batch signatures [180], but Stanek [262] showed that this method is flawed. Ferrara et al. [110] gave performance measurements for the schemes herein, and also showed how to batch verify other types of signatures, such as group and ring signatures. Law and Matt [178] pointed out that some identity-based signature schemes batch well, and gave methods for identifying invalid signatures in a batch. Because the goals of batch verification and aggregate signature are slightly different, it is important to clarify these distinct notions. Informally, batch verification's goal is to verify  $b$  distinct signatures quickly, while the goal in aggregate signatures is to compress



these signatures.

### 4.2.1 Definitions

We follow [38] for formalizing the aggregate signature in the chosen-key security model. For the convenience of our presentation, the attacker is also called an adversary or a forger.

**Definition 4.2.1.** The adversary  $\mathcal{A}$  is given a single public key, called challenge public key.  $\mathcal{A}$ 's goal is the existential forgery of an aggregate signature.  $\mathcal{A}$  is granted the power to choose all public keys except the challenge public key and access to a signing oracle on the challenge public key.  $\mathcal{A}$ 's advantage, denoted  $\text{Adv}_{\text{AggSig}, \mathcal{A}}$ , is defined to be the probability of success in the following game:

- **Setup.**  $\mathcal{A}$  is provided with a public key  $pk_1$ , generated at random.
- **Queries.**  $\mathcal{A}$  requests hash values of messages and signatures with  $pk_1$  on messages of his choice.
- **Response.** Finally,  $\mathcal{A}$  outputs  $b - 1$  additional public keys  $pk_2, \dots, pk_b$ . These keys, along with the initial key  $pk_1$ , will be included in  $\mathcal{A}$ 's forged aggregate.  $\mathcal{A}$  also outputs messages  $m_1, m_2, \dots, m_b$ , and an aggregate signature  $\sigma$ , where  $\sigma$  is signed by the  $b$  users and each user signs on the corresponding message.

The aggregate signature  $\sigma$  is called nontrivial if  $\mathcal{A}$  did not request a signature on  $m_1$  under  $pk_1$ . We define that  $\mathcal{A}$  wins if  $\sigma$  is nontrivial valid aggregate signature on messages  $m_1, m_2, \dots, m_b$  under keys  $pk_1, pk_2, \dots, pk_b$ . The probability is over the coin tosses of the key-generation algorithm and of  $\mathcal{A}$ .

**Definition 4.2.2.** An adversary  $\mathcal{A}$   $(t, q_H, q_S, b_{max}, \epsilon)$ -breaks an aggregate signature scheme in the chosen-key model if:  $\mathcal{A}$  runs in time at most  $t$ ;  $\mathcal{A}$  makes at most  $q_H$  queries to the hash function and at most  $q_S$  queries to the signing oracle;  $\text{Adv}_{\text{AggSig}, \mathcal{A}}$

is at least  $\epsilon$ ; and the forged aggregate signature is signed by at most  $b_{max}$  users. An aggregate signature scheme is  $(t, q_H, q_S, b_{max}, \epsilon)$ -secure against existential forgery in the chosen-key model if no forger  $(t, q_H, q_S, b_{max}, \epsilon)$ -breaks it.

Recall that a digital signature scheme is a tuple of algorithms (**Gen**, **Sign**, **Verify**) that also is correct and secure. The correctness property states that for all  $\mathbf{Gen}(1^\ell \rightarrow (pk; sk))$ , the algorithm  $\mathbf{Verify}(pk, m, \mathbf{Sign}(sk, m)) = 1$ .

**Definition 4.2.3** (Camenisch et al. [60]). Let  $\ell$  be the security parameter. Suppose  $(\mathbf{Gen}, \mathbf{Sign}, \mathbf{Verify})$  is a signature scheme and  $(pk_1, sk_1), \dots, (pk_b, sk_b)$  are independently generated according to  $\mathbf{Gen}(1^\ell)$ . We define a probabilistic batch verification algorithm (**Batch**) as follows:

- **Batch** $((pk_1, m_1, \sigma_1), \dots, (pk_b, m_b, \sigma_b)) = 1$  provided  $\mathbf{Verify}(pk_i, m_i, \sigma_i) = 1$  for all  $i = 1, 2, \dots, b$ .
- **Batch** $((pk_1, m_1, \sigma_1), \dots, (pk_b, m_b, \sigma_b)) = 0$  except with probability negligible in  $\epsilon$ , taken over the randomness of **Batch**, provided  $\mathbf{Verify}(pk_i, m_i, \sigma_i) = 0$  for any  $i = 1, 2, \dots, b$ .

**Definition 4.2.4** (Gentry and Ramzan [121]). The security model for identity based aggregate signature (IBAS) is defined as follows:

- **Setup:** The adversary  $\mathcal{A}$  is given public key  $pk$  of the Private Key Generation (**PKG**), an integer  $b_{max}$ , and any other needed parameters.
- **Queries:** Proceeding adaptively,  $\mathcal{A}$  may choose identities  $ID_i$  and request the private key  $sk_i$ . Also,  $\mathcal{A}$  may request an IBAS  $\sigma_S$  on  $(pk, S, \{m_i\}_{i=1}^{k-1})$  where  $S = \{ID_i\}_{i=1}^{k-1}$ . We require that  $\mathcal{A}$  has not to make a query  $(pk, S', \{m_i\}_{i=1}^{k-1})$  where  $ID_i \in S \cap S'$  and  $m'_i \neq m_i$ .
- **Response:** For some  $(pk, \{ID_i\}_{i=1}^b, \{m_i\}_{i=1}^b)$  with  $b \leq b_{max}$ ,  $\mathcal{A}$  outputs an IBAS  $\sigma_l$ .

We call that the signature is nontrivial, where for some  $i$ ,  $1 \leq i \leq b$ ,  $\mathcal{A}$  did not request the private key for  $ID_i$  and did not request a signature including the pair  $(ID_i, m_i)$ . We define that  $\mathcal{A}$  wins if  $\sigma_b$  is a nontrivial valid signature on  $(pk, \{ID_i\}_{i=1}^b, \{m_i\}_{i=1}^b)$

**Definition 4.2.5.** An IBAS adversary  $\mathcal{A}$   $(t, \epsilon, b_{max}, q_H, q_E, q_S)$ -breaks an IBAS scheme in the above model if: for the integer  $b_{max}$  as above,  $\mathcal{A}$  runs in time at most  $t$ ;  $\mathcal{A}$  makes at most  $q_H$  hash function queries,  $q_E$  private key extraction queries, and at most  $q_S$  signing oracle queries; and  $\text{Adv}_{\text{IBAS}, \mathcal{A}}$  is at least  $\epsilon$ . An IBAS scheme is  $(t, \epsilon, b_{max}, q_H, q_E, q_S)$ -secured against existential forgery if no adversary  $(t, \epsilon, b_{max}, q_H, q_E, q_S)$ -breaks it.

## 4.2.2 Aggregate Signature

Aggregation is obtained by combining the resulting  $b$  signatures  $\{\sigma_1, \sigma_2, \dots, \sigma_b\}$  into one aggregate signature  $\sigma$ :

$$\sigma \leftarrow \prod_{i=1}^b \sigma_i.$$

We introduce aggregate signature scheme given by Boneh et al.'s [38] as follows:

- **Setup:** Let  $e : \mathbb{G}_1 \times \mathbb{G}_2 \rightarrow \mathbb{G}_T$ , where  $g$  generates the group  $\mathbb{G}_2$  of prime order  $q$ .
- **Gen:** Chooses a random  $sk \in \mathbb{Z}_q$  and outputs  $pk = g^{sk}$ .
- **Sig:** A signature on message  $m$  is  $\sigma = H(m)^{sk}$ , where  $H$  is a hash function  $H : \{0, 1\}^* \rightarrow \mathbb{G}_1$ .
- **Ver:** To verify signature  $\sigma$  on message  $m$ , one checks that  $e(\sigma, g) \stackrel{?}{=} e(H(m), pk)$ .

Batch verification is very simple. What we need to do are the following two steps: checking that the  $m_i$ 's are mutually distinct and ensuring that

$$e(g, \sigma) = \prod_{i=1}^b e(pk_i, h_i) \quad (4.1)$$

where  $h_i = h(m_i)$ . Equality (4.1) holds because

$$e(g, \sigma) = e(g, \prod_{i=1}^b h_i^{sk_i}) = \prod_{i=1}^b e(g, h_i)^{sk_i} = \prod_{i=1}^b e(g^{sk_i}, h_i) = \prod_{i=1}^b e(pk_i, h_i). \quad (4.2)$$

In the special case when all  $b$  signatures are issued by the same public key  $pk_0$  (same signer), the verification is faster *only for 2 pairings*. One needs to verify that

$$e(g, \sigma) = e(\prod_{i=1}^b h(m_i), pk_0),$$

where  $m_1, m_2, \dots, m_b$  are the signed messages.

However, Camenisch et al. [60] pointed out that the aggregate scheme in [38] is not a batch verification scheme since, for any  $a \neq 1 \in \mathbb{G}_1$ , the two invalid message-signature pairs  $P_1 = (m_1, a \cdot H(m_1)^{sk})$  and  $P_2 = (m_2, a^{-1} \cdot H(m_2)^{sk})$  will verify under Definition 4.2.3.

### 4.2.3 Identity-Based Aggregate Signature

In 2006, Gentry and Ramzan [121] gave the first identity-based aggregate signature (IBAS) as follows:

- **Setup:** It generates groups  $\mathbb{G}_1$  and  $\mathbb{G}_2$  of prime order  $q$  and an admissible pairing  $e : \mathbb{G}_1 \times \mathbb{G}_1 \rightarrow \mathbb{G}_2$ ; chooses an arbitrary generator  $P \in \mathbb{G}_1$ . In addition, it chooses three cryptographic hash functions  $H_1, H_2 : \{0, 1\}^* \rightarrow \mathbb{G}_1$ , and  $H_3 : \{0, 1\}^* \rightarrow \mathbb{Z}_q^*$ . After setting the system parameters, **PKG** picks a random number

$sk_{PKG} \in \mathbb{Z}_q^*$  as its secret key and sets its public key as  $P_{pub} = sk_{PKG} \cdot P$ .

The system parameters are  $params = (\mathbb{G}_1, \mathbb{G}_2, q, e, P, P_{pub}, H_1, H_2, H_3)$ . The system's secret is  $sk_{PKG}$ .

• **Extract:**

$$sk_i^0 = sk_{PKG} \cdot Q_{ID_i}^0 \text{ and } sk_i^1 = sk_{PKG} \cdot Q_{ID_i}^1$$

where  $Q_{ID_i}^0 = H_1(ID_i \| 0)$  and  $Q_{ID_i}^1 = H_1(ID_i \| 1)$ .

• **Individual Signing:**

- (1) Choose a unique string  $\omega$  and compute  $P_\omega = H_2(\omega) \in \mathbb{G}_1$ ;
- (2) Compute  $c_i = H_3(m_i, ID_i, \omega) \in \mathbb{Z}_q^*$ ;
- (3) Generate a random number  $r_i$ , where  $r_i \in \mathbb{Z}_q^*$ ;
- (4) Compute signature  $Sig_i = (\omega, U_i, V_i)$  for message  $m_i$ , where

$$U_i = r_i \cdot P \text{ and } V_i = r_i \cdot P_\omega + sk_i^0 + c_i \cdot sk_i^1. \quad (4.3)$$

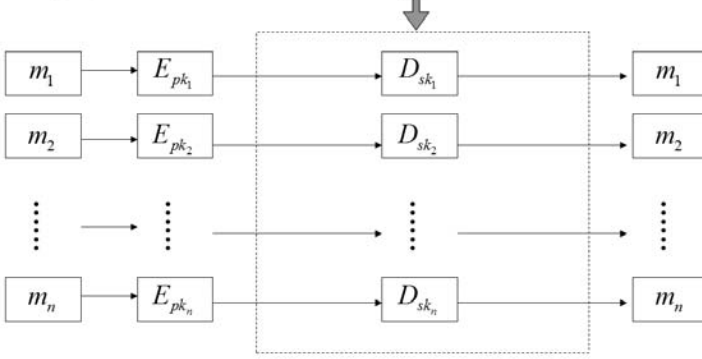
- **Aggregate:** Anyone can aggregate a collection of individual signatures that use the same string  $\omega$ . Signatures  $Sig_i = (U_i, V_i)$  where  $0 \leq i \leq N$  can be aggregated to  $Sig = (U, V)$  where  $U = \sum_{i=0}^b U_i$  and  $V = \sum_{i=0}^b V_i$ .

• **Verification:**

$$e(V, P) = e(P_\omega, U) e\left(\sum_{i=1}^b Q_{ID_i}^0 + \sum_{i=1}^b c_i \cdot Q_{ID_i}^1, P_{pub}\right) \quad (4.4)$$

## 通用的加密模型

decrypt operation cost:  $n$  time individual decryption



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Figure 4.1: Individual decryption model

where  $c_i = H_3(m_i, ID_i, \omega)$ . The batch verification equality (4.4) holds because:

$$\begin{aligned}
 e(V, P) &= e\left(\sum_{i=1}^b V_i, P\right) = e\left(\sum_{i=1}^b r_i P_\omega + \sum_{i=1}^b sk_i^0 + \sum_{i=1}^b c_i sk_i^1, P\right) \\
 &= \prod_{i=0}^i e(r_i P_\omega, P) \prod_{i=0}^i e(sk_{PKG} Q_{ID_i}^0, P) \prod_{i=0}^i e(c_i sk_{PKG} Q_{ID_i}^1, P) \\
 &= \prod_{i=0}^i e(P_\omega, r_i P) \prod_{i=0}^i e(Q_{ID_i}^0, sk_{PKG} P) \prod_{i=0}^i e(c_i Q_{ID_i}^1, sk_{PKG} P) \\
 &= e\left(P_\omega, \sum_{i=1}^b U_i\right) e\left(\sum_{i=1}^b Q_{ID_i}^0, P_{pub}\right) e\left(\sum_{i=1}^b c_i Q_{ID_i}^1, P_{pub}\right) \\
 &= e(P_\omega, U) e\left(\sum_{i=1}^b Q_{ID_i}^0 + \sum_{i=1}^b c_i Q_{ID_i}^1, P_{pub}\right)
 \end{aligned}$$

### 4.3 Batch Decryption and Batch Key Agreement

In the above sections, we focus on signature aggregation and batch verification. From now on, we focus our attention to batch decryption and key agreement schemes. In batch decryption, batch RSA based on the RSA scheme is a frequently used example to explain this concept. Batch RSA was first proposed by Fiat [111] in 1989 and later developed by Boneh and Shacham [43] in 2002. It is designed to speed up RSA decryp-

## 通用的加密模型

decrypt operation cost: at the same level of every  $D_{sk_i}$

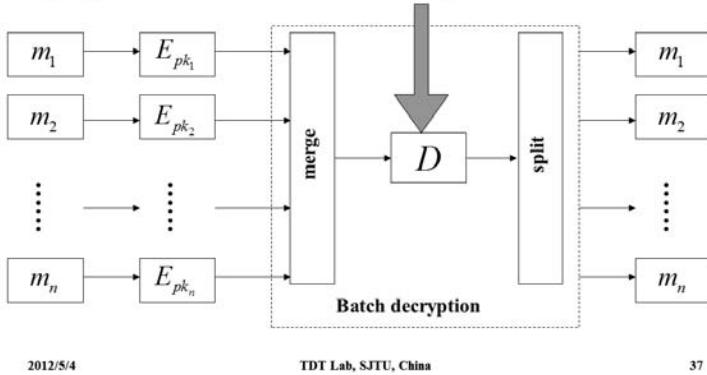


Figure 4.2: Universal batch decryption model

tion and signing algorithms. Batch RSA has many applications. Since it was proposed, it has attracted many researchers and industry engineers. Shacham and Boneh [242] improved performance of SSL based on batch RSA. Yacobi and Beller [298] applied the mathematical ideas of batch RSA to DH-based schemes. They proposed a batch Diffie–Hellman key agreement scheme. Figure 4.1 gives an individual decryption model where messages are encrypted by the different public keys and sent to decryption module individually and the decryption operations are performed one by one. Thus, the total cost is  $n$  times individual decryption in the dotted line box (we assume that it is a server) in Figure 4.1.

We are trying to find a way to improve the efficiency by a batch model instead of this individual decryption model. Figure 4.2 gives a universal batch decryption model where the messages are encrypted by the different public keys and sent to decryption module individually. The decryption module first merges these ciphertexts, decrypts using its secret key *only once*, and splits to individual plaintexts.

Batch RSA can do a number of RSA decryptions for approximately the cost of a single decryption.

For the integrally introducing batch RSA, we first give the RSA scheme and then give Fiat’s batch RSA and Boneh and Shacham’s batch RSA.

### 4.3.1 Review of RSA

RSA [233] is the most widely deployed public key cryptosystem. It is used for securing web traffic, e-mail, and some wireless devices. Since RSA is based on arithmetic modulo large integer numbers it might be slow in constrained environments. For example, 1024-bit RSA decryption on a small handheld device such as the PalmPilot III can take as long as 30 seconds. Similarly, on a heavily loaded web server, RSA decryption significantly reduces the number of SSL requests per second that the server can handle. Typically, one improves performance of RSA using special-purpose hardware. Current RSA coprocessors can perform as many as 10,000 RSA decryptions per second (using a 1024-bit modulus) and even faster processors are coming out.

The RSA encryption scheme can be summarized as follows. The public key is the pair  $(N, e)$ , where  $N$  is the product of two large primes  $p$  and  $q$ , and  $e$  is chosen to be coprime to Euler's function  $\phi(N)$ . The private key is  $d$  where  $d \equiv e^{-1} \pmod{\phi(N)}$ .

To encrypt a message  $m$ , one computes the ciphertext  $c$  as follows:

$$c \equiv m^e \pmod{N} \quad (4.5)$$

To decrypt the ciphertext  $c$ , one computes

$$m \equiv c^d \pmod{N} \equiv c^{1/e} \pmod{N}. \quad (4.6)$$

Thus, every decryption consists of one full-sized modular exponentiation.

Later, Quisquater and Couvreur [230] suggested the use of the Chinese Remainder Theorem (CRT) in order to make the decryption run slightly faster. It is standard practice to employ the Chinese Remainder Theorem for RSA decryption. Rather than compute  $\sigma \leftarrow c^d \pmod{N}$ , the decryptor evaluates:

$$m_p \leftarrow c_p^{d_p} \pmod{p} \quad (4.7)$$



and

$$m_q \leftarrow c_q^{d_q} \pmod q \quad (4.8)$$

Here,  $d_p \equiv d \pmod{p-1}$  and  $d_q \equiv d \pmod{q-1}$ . Then the decryptor uses the CRT to calculate  $m$  from  $m_p$  and  $m_q$ . This is approximately four times faster than evaluating  $c^d \pmod N$  directly.

Generally speaking, a small value of  $e$  is chosen, which is coprime to  $\phi(N)$ , and  $d$  will be  $\mathcal{O}(\phi(N))$ . In fact, if  $d$  was too small (less than exponential in the security parameter), it would allow the cryptanalyst to attack the scheme. Wiener [296] gives attacks on the short private key  $d$  of RSA.

*Problem 4.3.1.* How to cut down the cost of decryption with a batch of ciphertext?

Fiat's batch RSA gives us the first answer to this problem.

### 4.3.2 Batch RSA

Fiat's batch RSA [111] can be shown in Figure 4.3. Let  $e_1, e_2, \dots, e_b$  be  $b$  different

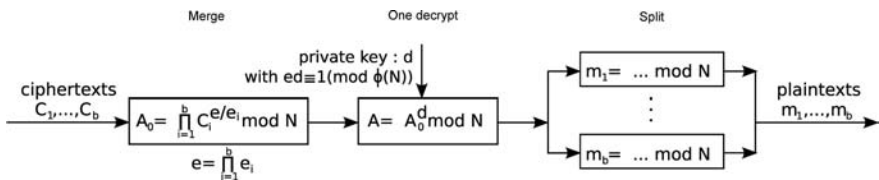


Figure 4.3: Batch RSA decryption

encryption exponents, coprime to  $\phi(N)$  and to each other where  $N$  is defined as in the RSA scheme. Let

$$e = \prod_{i=1}^b e_i \quad (4.9)$$

be  $\mathcal{O}$  bits long. Given ciphertext  $c_1, c_2, \dots, c_b$  related to plaintext  $m_1, m_2, \dots, m_b$ , our goal is to generate the  $b$  roots (decryptions):

$$c_1^{1/e_1}, c_2^{1/e_2}, \dots, c_b^{1/e_b} \pmod{N}. \quad (4.10)$$

Let  $T$  be a binary tree with leaves labeled  $e_1, e_2, \dots, e_b$ . Let  $d_i$  denote the depth of the leaf labeled  $e_i$ ,  $T$  should be constructed so that  $W = \sum_{i=1}^b d_i \log e_i$  is minimized similar to the Huffman code tree construction. For the main result of  $\mathcal{O}(\log^2 N)$  multiplications per RSA operation, Fiat's scheme could simply assume that  $T$  is a full binary tree. In practice, there is some advantage in using a tree that minimizes the sum of weight times path length because the work performed is proportional to the sum  $W$  above. Note that  $W = \mathcal{O}(\log b \log e)$ . Finally, Fiat shows that the number of multiplications required to compute the  $b$  roots above is  $\mathcal{O}(W + \log N)$ .

In Fiat's scheme, the first goal is to generate the product:

$$A_0 = c_1^{e/e_1} \cdot c_2^{e/e_2} \cdot \dots \cdot c_b^{e/e_b}. \quad (4.11)$$

Fiat also shows that this requires  $\mathcal{O}(W)$  multiplications. Use the binary tree  $T$  as a guide, working from the leaves to the root. Every internal node takes the recursive result from the left branch ( $L$ ) raises it to the power  $E_R$  where  $E_R$  is the product of the labels associated with leaves on the right branch. Similarly, each node takes the result from the right branch ( $R$ ) and raises it to the power  $E_L$ , which is the product of the labels on the left branch. Each node saves the intermediate results  $L^{E_R}$  and  $R^{E_L}$  (required later). The result associated with this node is  $L^{E_R} \cdot R^{E_L}$ . The product  $A_0$  is simply the result associated with the root.

Second, it extracts the  $e$ -th root of the product  $A_0$ :

$$A = A_0^{1/e} = c_1^{1/e_1} \cdot c_2^{1/e_2} \cdot \dots \cdot c_b^{1/e_b} \pmod{N} \quad (4.12)$$

This involves  $\mathcal{O}(\log N)$  modular multiplications which is equivalent to one RSA decryption.

Third, the factors of  $A$  are the roots. The next goal is to break the product  $A$  into two subproducts, the breakup is implied by the structure of the binary tree  $T$  used to generate the product  $C$ . The scheme repeats this recursively to break up the product into its  $b$  factors.

Let  $e_1, e_2, \dots, e_k$  be the labels associated with the left branch of the root of the binary tree  $T$ . Define an exponent  $X$  by means of the Chinese Remainder Theorem (CRT):

$$\left\{ \begin{array}{l} X \equiv 0 \pmod{e_1} \\ X \equiv 0 \pmod{e_2} \\ \dots \\ X \equiv 0 \pmod{e_k} \\ X \equiv 1 \pmod{e_{k+1}} \\ \dots \\ X \equiv 1 \pmod{e_b} \end{array} \right. \quad (4.13)$$

There is a unique solution for  $X$  modulo  $e = \prod_{i=1}^b e_i$  by CRT. By (4.13), let

$$X = \left( \prod_{i=1}^k e_i \right) \cdot X_1 \quad (4.14)$$

and

$$X - 1 = \left( \prod_{i=k+1}^b e_i \right) \cdot X_2, \quad (4.15)$$

where  $X_1, X_2$  are positive integers. Let  $P_1 = \prod_{i=1}^k e_i$  and  $P_2 = \prod_{i=k+1}^b e_i$ , then  $X = P_1 \cdot X_1$  and  $X - 1 = P_2 \cdot X_2$ . Compute

$$C_1 = c_1^{P_1/e_1} \cdot c_2^{P_1/e_2} \cdot \dots \cdot c_k^{P_1/e_k} \pmod{N} \quad (4.16)$$

and

$$C_2 = c_{k+1}^{P_2/e_{k+1}} \cdot c_{k+2}^{P_2/e_{k+2}} \cdot \dots \cdot c_b^{P_2/e_b} \pmod{N} \quad (4.17)$$

Note that  $C_1$  and  $C_2$  have been computed, as the left and right branch values of the root, during the tree-based computation of  $A_0$ .

Raise  $A$  to the  $X$ -th power modulo  $N$ , it follows from

$$\begin{aligned} A^X &= \left( \prod_{i=1}^b c_i^{1/e_i} \right)^X \\ &= \left( \prod_{i=1}^k c_i^{1/e_i} \right)^{P_1 X_1} \cdot \left( \prod_{i=k+1}^b c_i^{1/e_i} \right)^{P_2 X_2} \cdot \prod_{i=k+1}^b c_i^{1/e_i} \\ &= C_1^{X_1} \cdot C_2^{X_2} \cdot \prod_{i=k+1}^b c_i^{1/e_i}, \end{aligned}$$

that  $Q_2 = \prod_{i=k+1}^b c_i^{1/e_i}$ , and so  $Q_1 = \prod_{i=1}^k c_i^{1/e_i}$  since  $Q_1 = A/Q_2$ . After that, it recursively breaks up the two products  $Q_1$  and  $Q_2$ , respectively, until each  $c_i^{1/e_i}$  is obtained. Thus, the overall number of multiplications is  $\mathcal{O}(W)$ . The number of modular divisions required is  $\mathcal{O}(b)$ .

To summarize this method, we introduce the following theorem.

**Theorem 4.3.2** (Fiat [111]). *Let  $e_1, e_2, \dots, e_b$  be  $b$  different encryption exponents, relatively prime to  $\phi(N)$  and to each other. Given the ciphertexts  $c_1, c_2, \dots, c_b$ , we can generate the  $b$  roots*

$$c_1^{1/e_1}, c_2^{1/e_2}, \dots, c_b^{1/e_b} \pmod{N}$$

*in  $\mathcal{O}(\log b(\sum_{i=1}^b \log e_i) + \log N)$  modular multiplications and  $\mathcal{O}(b)$  modular divisions.*

Fiat's batch RSA gave an example when using small public exponents  $e_1$  and  $e_2$ , it decrypts two ciphertexts for approximately the price of one decryption.

Suppose  $c_1$  is a ciphertext obtained by encrypting some plaintext  $m_1$  using the public key  $(N, 3)$ , and  $c_2$  is a ciphertext for some  $m_2$  using  $(N, 5)$ . To decrypt, they must

compute  $c_1^{1/3}$  and  $c_2^{1/5} \pmod{N}$ . Fiat observed that by setting

$$A = (c_1^5 \cdot c_2^3)^{1/15}, \quad (4.18)$$

one obtains

$$c_1^{1/3} = \frac{A^{10}}{c_1^3 \cdot c_2^2} \quad (4.19)$$

and

$$c_2^{1/5} = \frac{A^6}{c_1^2 \cdot c_2}. \quad (4.20)$$

Hence, at the cost of computing a single 15th root and some additional arithmetic, it is able to decrypt both  $c_1$  and  $c_2$ . Computing a 15th root is considered to take the same time as a single RSA decryption in this case.

Boneh and Shacham [43] generalized (4.19) and (4.20) to

$$m_i \equiv c_i^{1/e_i} \equiv \frac{A^{\alpha_i}}{c_i^{(\alpha_i-1)/e_i} \cdot \prod_{j \neq i} c_j^{\alpha_i/e_j}} \pmod{N} \quad (4.21)$$

where  $\alpha_i$  satisfies

$$\alpha_i \equiv \begin{cases} 1 & \pmod{e_i}, \\ 0 & \pmod{e_j}, \quad \text{for } j \neq i, \end{cases} \quad (4.22)$$

where  $1 \leq i, j \leq b$ .

This method requires  $b$  different modular inversions whereas Fiat's tree-based method requires  $2b$  modular inversions, but fewer auxiliary multiplications. Thus, Boneh and Shacham give an improvement to Fiat's scheme. However, in Boneh and Shacham's scheme, it is still required to solve the systems of linear congruences Equations (4.22) with  $b$  modular inversions. In addition, it is not easy to implement in hardware using these different public keys  $(e_1, e_2, \dots, e_b)$  as the modular equipped in advance in the chip based systems. Thus, we leave the following problems:

*Problem 4.3.3.* For each  $b$ , how to obtain all  $\alpha_i$ 's satisfying Equation (4.22) without any modular operation?

*Problem 4.3.4.* Are there any other methods to split the plaintext directly like Equation (4.21)?

Once Problem 4.3.3 is solved, we can directly obtain  $m_i$  satisfying Equation (4.21) using only one modular  $N$ , which allows us to implement batch RSA in hardware easily. If the answer to Problem 4.3.4 is positive, we can get more methods to split plaintext to individual. We will discuss it in Section 4.4.

### 4.3.3 Batch Key Agreement

In this section, we consider a star-like network to present the batch key agreement. For example, in a personal communication system network, the central server, through the ports, agrees upon a session key with each of the users. Each port holds a secret key that is used to establish a secret channel in the network.

Yacobi and Beller [298] proposed a solution, in which each of the users picks a random number, encrypts it with the public key of the corresponding port, and sends it to the counterpart port. The central server decrypts the received ciphertext, and uses the random numbers or their hash values as session keys. Yacobi and Beller [298] summarized the difference between RSA-based and DH-based key exchange. The basic DH scheme was first proposed in [99], which makes the key exchanged individually and needs  $b$  operations when the number of the users is  $b$ . Yacobi and Beller [298] presented a batch RSA key agreement as follows:

- First, the central server with  $b$  ports communicates with the users. The central server generates a public/private key pair for each port. The modulus is a composite  $N$ , which is the product of two large primes  $p$  and  $q$ . Each port's public key  $e_i$  is chosen to be prime to Euler's function  $\phi(N)$  and pairwise coprime. The private key is  $d_i$ , where  $d_i \equiv e_i^{-1} \pmod{\phi(N)}$ .

- Second, when  $b$  users want to agree on session keys individually with the central server, each user needs to find an available port and encrypts a chosen random number using the port's public key  $e_i$ .
- Third, when the central server receives the ciphertexts from each port  $i$ , it takes one decryption similar to the batch RSA decryption in the above section and gets the random numbers that could finally generate the session keys.

## 4.4 Batch RSA's Implementation Based on Diophantine Equations

In this section, we introduce our batch RSA implementation based on Diophantine Equations. Our implementation of batch RSA reduces  $b - 1$  modular inversion computation compared with Fiat's batch RSA and Boneh and Shacham's implementations. Furthermore, our implementation is more suitable for hardware circuits. If we apply this algorithm to a hardware circuit for batch RSA decryption, our implementation requires only one circuit module to do modular operations for  $N$ , while each of the other two implementations requires  $b$  different modules to do modular operations, which is difficult to be written into chips in advance and varies for every decryption.

### 4.4.1 Implementation Based on Plus-Type Equations

We find that the solutions of the Diophantine equation

$$\sum_{1 \leq i \leq n} \frac{1}{x_i} + \prod_{1 \leq i \leq n} \frac{1}{x_i} = 1, \quad 2 \leq x_1 < x_2 < \cdots < x_n \quad (4.23)$$

can be used to implement batch RSA so that the plaintexts can be constructed directly and the whole decryption only uses modular  $N$ . We call Equation (4.23) a *Plus-Type Diophantine Equation* or *Plus-Type Equation*. For convenience, throughout this book,

Equation (4.23) with  $n$  unknown variables is denoted as Equation (4.23) <sub>$n$</sub> .

Our new implementation of batch RSA contains three algorithms: KeyGen, Encrypt, and Decrypt. The essence of the new implementation is in the decryption algorithm.

- **KeyGen:** This algorithm follows Fiat's batch RSA key generation algorithm except choosing the public keys  $e_1, e_2, \dots, e_b$  from solutions of Equation (4.23) <sub>$n$</sub> , that is,  $\{e_1, e_2, \dots, e_b\} \subseteq \{x_1, x_2, \dots, x_n\}$ .
- **Encrypt:** This algorithm is same as in Fiat's batch RSA.
- **Decrypt:** This algorithm consists of three steps:
  - “Merge”: The algorithm evaluates  $A_0 = \prod_{1 \leq i \leq b} c_i^{e/e_i} \pmod N$ . This computation can be finished by Fiat's optimized algorithm [111].
  - “One decrypt”: The algorithm evaluates one full-scale modular exponentiation  $A \equiv A_0^d \pmod N$ .
  - “Split”: The essence of our implementation is

$$\alpha_i = \frac{\prod_{1 \leq j \leq n} x_j}{e_i}, \quad (4.24)$$

where  $1 \leq i \leq b$ , and it outputs<sup>1</sup>

$$m_i \equiv \frac{c_i^{(\alpha_i+1)/e_i} \cdot \prod_{j \neq i} c_j^{\alpha_i/e_j}}{A^{\alpha_i}} \pmod N. \quad (4.25)$$

*Remark 4.4.1.* In our decrypt algorithm, after the plaintext  $(m_1, m_2, \dots, m_{b-1})$  are decrypted, the final  $m_b$  can be calculated by

$$m_b \equiv \frac{A}{\prod_{j=1}^{b-1} m_j} \pmod N, \quad (4.26)$$

---

<sup>1</sup>Equation (4.25) is slightly different from Equation (4.22) since we choose the Plus-Type Equation.



where the computation can be reduced.

**Remark 4.4.2.** If the public keys  $e_1, e_2, \dots, e_b$  are chosen arbitrarily (not from the solution  $(x_1, x_2, \dots, x_n)$  of Equation (4.23)<sub>n</sub>), Equation (4.25) also satisfies when  $\alpha_i$  satisfies

$$\alpha_i \equiv \begin{cases} -1 & (\text{mod } e_i), \\ 0 & (\text{mod } e_j), \quad \text{for } j \neq i. \end{cases} \quad (4.27)$$

**Theorem 4.4.3** (Correctness). *The above scheme is correct.*

*Proof.* From  $A \equiv A_0^d \pmod{N}$ , we have

$$A \equiv \prod_{1 \leq i \leq b} c_i^{1/e_i} \pmod{N}.$$

Since  $(x_1, \dots, x_n)$  is a solution of Equation (4.23)<sub>n</sub>, we have

$$\frac{\prod_{1 \leq k \leq n} x_k}{x_i} \equiv -1 \pmod{x_i}$$

and  $\frac{\prod_{1 \leq k \leq n} x_k}{x_i} \equiv 0 \pmod{x_j}$  (for all  $i, j$  such that  $1 \leq i \leq n, 1 \leq j \leq n, j \neq i$ ).

Hence  $e_i | \alpha_i + 1$  and  $e_j | \alpha_i$  (for all  $i, j$  such that  $1 \leq i \leq b, 1 \leq j \leq b, j \neq i$ ). In Equation (4.25), all the modular exponentiations are integer exponentiations. Therefore, it follows from a direct computation that  $m_i \equiv c_i^{1/e_i} \pmod{N}$ , where  $1 \leq i \leq n$ .  $\square$

For a simple example, we choose a solution  $(2, 3, 7)$  of Equation (4.23)<sub>3</sub> and take 3 and 7 as the public keys  $e_1$  and  $e_2$ , respectively. Thus,

$$A = (c_1^7 \cdot c_2^3)^{1/21}, \quad (4.28)$$

$$\alpha_1 = \frac{\prod_{1 \leq j \leq 3} x_j}{e_1} = \frac{2 \cdot 3 \cdot 7}{3} = 14, \quad (4.29)$$

and

$$\alpha_2 = \frac{\prod_{1 \leq j \leq 3} x_j}{e_2} = \frac{2 \cdot 3 \cdot 7}{7} = 6. \quad (4.30)$$

Thus, we can get  $m_i$  as follows:

$$m_1 \equiv \frac{c_1^5 \cdot c_2^2}{A^{14}} \pmod{N}, \quad (4.31)$$

and

$$m_2 \equiv \frac{c_1^2 \cdot c_2}{A^6} \pmod{N}. \quad (4.32)$$

In fact,  $m_b$  can be obtained through Remark 4.4.1. As shown in the above example, if we compute  $m_2$  first and then compute  $m_1 = \frac{A}{m_2}$ , we can finish the split (decryption) efficiently. Therefore, we give a new method to compute  $\alpha_i$  without any modular operation and a new method to split the plaintext. And so, the answer to Problem 4.3.4 is “YES”. As far as Problem 4.3.3 is concerned, for arbitrary positive integer  $b$ , we are able to find many  $\alpha_i$ 's. Furthermore, if  $b$  is not greater than 8, we are able to find all  $\alpha_i$ 's. We will introduce our work towards these problems in detail in Section 4.5.

Here, we would like to point out the differences among Fiat's RSA, Boneh and Shacham's Batch RSA and our implementation. The advantage of our implementation is that there is only one modular,  $\pmod{N}$ , while both Fiat's batch RSA and Boneh and Shacham's implementation require at least  $b$  modulars.

**Fiat's batch RSA:** In KeyGen,  $\{e_1, \dots, e_b\}$  are randomly picked. In Decrypt, Fiat's tree-based method requires  $2b + 1$  modulars.  $k$  is chosen so that  $1 \leq k < b$ . Let  $P_1 = \prod_{1 \leq i \leq k} e_i$ ,  $P_2 = \prod_{k < i \leq b} e_i$ ,  $X_1 = P_1^{-1} \pmod{P_2}$ ,  $X_2 = \frac{P_1 X_1 - 1}{P_2}$ . Then compute

$$\frac{A^{P_1 X_1}}{(\prod_{1 \leq i \leq k} c_i^{X_1/e_i})^{P_1} (\prod_{k < i \leq b} c_i^{X_2/e_i})^{P_2}} \pmod{N} = \prod_{k < i \leq b} m_i \pmod{N},$$

and

$$\frac{A}{\prod_{k < i \leq b} m_i} \pmod{N} = \prod_{1 \leq i \leq k} m_i \pmod{N}.$$

This process is done recursively until  $m_1, m_2, \dots, m_b$  are calculated.

**Boneh and Shacham's improvement:** In Decrypt,  $\alpha_i$  is obtained from Equations (4.22). Then  $m_1, \dots, m_{b-1}$  are evaluated by Equation (4.21).  $m_b$  is calculated by  $m_b = \frac{A}{\prod_{1 \leq j < b} m_j} \pmod{N}$ . The evaluation of  $\alpha_1, \alpha_2, \dots, \alpha_b$  requires  $b$  modulars.

**Our implementation:**  $\alpha_1, \alpha_2, \dots, \alpha_b$  are evaluated directly from the solution of Equation (4.23). Therefore,  $b - 1$  modulars are reduced. Furthermore, we get the plaintext  $m_i$  from Equation (4.21).

#### 4.4.2 A Concrete Example Based on Plus-Type Equations

We give a concrete example to explain our algorithm to show its advantage.

**KeyGen:** Let the batch size  $b = 7$ . Randomly choose two 512-bit primes  $p$  and  $q$ :

$p = 10117548898104292384286149905619010120991278203895691114024386173114298$   
 $44196706850901843445909220802938862195711111823138805089269233175589739$   
 $8024486601967,$

$q = 11942104782310447106884510125435655207765533763318483151065949254099976$   
 $07148379656268155866005262065749842699695077677106764786988072219334757$   
 $0651658382099.$

Let  $N = p \cdot q$ .

$N = 12082482908131106460128885334610470733722100002825862245001246171456821$   
 $80012763601040492827413417501769408868451698340800241737759849184798713$   
 $84311467341679006113447482380675139998093264657381057047502767534877118$   
 $98885364204757448057797883429393903920166667955622009797159159329896972$   
 $3360357448012917410988733.$

We also randomly choose a solution  $(x_1, x_2, \dots, x_8)$  of Equation (4.23)<sub>8</sub> where

$$(x_1, x_2, \dots, x_8) = (2, 5, 7, 11, 17, 157, 961, 4398619).$$

Choose the public keys as a subset of  $\{x_1, x_2, \dots, x_8\}$ , for example,

$$e_1 = 5, e_2 = 7, e_3 = 11, e_4 = 17, e_5 = 157, e_6 = 961, e_7 = 4398619.$$

We compute

$$x = \prod_{1 \leq i \leq 8} x_i = 8687184244716670.$$

Thus, the public key is  $(N, b = 7, e_1, \dots, e_7, x)$ . We compute the private key  $d \equiv (\prod_{1 \leq i \leq 7} e_i)^{-1} \pmod{\phi(N)}$ .  $d$  is as follows:

$d = 51876628765295878339434317579819334450607160465137205859602663251881894$   
 $76248150332865697383721278328392362980310483479730993621055438208436678$   
 $79229731524489835689222861536030650222353925195560713659815910775865522$   
 $82418304145301163771498370785315888078449179919532145099665908400332933$   
 $900700629001536377165451.$

**Encrypt:** In this algorithm, we randomly choose a set of plaintexts as follows:

$m_1 = 42216335483922430422828646305226638636955551187036758338218940369050771$   
 $71838777670539200939753400187558122765253611872592118414072653271581340$   
 $49557150686378816189760567012318705743611765112882799638223233870402496$   
 $60612305070655416046124619992356068472712803701841057530336101962637429$   
 $708518613124606072032913,$

$m_2 = 11515339478373653799409100162804092001627376847524519454456897879722755$   
 $72450525649782084136951971592156312409505014881084133776507758249194091$   
 $01745183559536347953424619444068351729672938002818082476781228433719588$   
 $49030458058556915325246572500148818313879322593916775020883447348607809$   
 $3654331805097407477851212,$

$m_3 = 47564196149813970007363884131737190884763138301794058772403642723080473$   
 $12035300434644530039033892729163849746746554561004845842905132968056260$   
 $46689776922999970386770238498695474858361140251559169316946590930490790$   
 $74855077783622152296547567667859890106605979365108882896772768642533878$   
 $129671533527391835259813,$

$m_4 = 47076113695544237253122064671136422055958210860752623025678325581081815$   
 $84726955718832078797363240141336812574782622687307406041247418256240938$   
 $54400992362572138986858352204679423469820466914319036593211022030949477$   
 $53860823597327978844407313209862946613671724072935020256407939759987354$   
 $078957243653508906462050,$

$m_5 = 30568044214000539098104805334573980758657958880378082233476394264866596$   
 $11473267935711445240332623905770776659086150428617242444099500783232312$   
 $04676910816693107200514851741175379334306919644777051154134292280179732$   
 $07188602119552456911467534569434906151478718718595432509630542457525910$   
 $131946213578956292963396,$

$m_6 = 41377747175811034232018876246074059499644813172580074927939214929882998$   
 $79343270705904474700195302890675803041195770969714258620664937410522441$   
 $78736605455167725304146829924000744594614850114240963709159146374669699$   
 $62099112312462790622604288246922576918641763721175549148057922321723822$   
 $247230264707138113937042,$

and

$m_7 = 18771896161902354107492151420608960655175748847557766591096333115544693$   
 $88639115772459886403263920691549973938090162854541237354779139974823373$   
 $22150711283782344174258728142522160188013242155927394712955764457748723$   
 $37172809149112059845297173025509381745188281479782761341318653942719134$   
 $671008123932040544891767.$

Since the ciphertext  $c_i$  is computed as follows  $c_i \equiv m_i^{e_i} \pmod{N}$ , where  $1 \leq i \leq 7$ .

The ciphertexts are

$c_1 = 32162445244401743657170446510623926600937713705351525278004836278347467$   
 $78927818565931306900626346594961086912725141809256522635855273218400519$   
 $72284664046732942258138795733271680027669332436406173493987951365152414$   
 $60797459530585405018120324737308623849797695612736566101235396092740354$   
 $722520913907481538932914,$

$c_2 = 11126467052899354989487810154732066615105697509420849586673823566312709$   
 $37106371291521967449468618459618164119164667029310985402655564125542964$   
 $94010110454544233209099018186584811824903310123116178433727720898855237$   
 $28984076826348484900130452142401788332892321945373838606441449511625255$   
 $9429226960199357541705876,$

$c_3 = 42702738253874326011706347202093584148640737529891335287560057328232353$   
 $12936031281751279979434019357164808476168319559847561796277935971011335$   
 $12500689235545118116809962307591329406164127984458850739739240726187291$   
 $08244500104462787099197514522083354004527717032842340463375898989280266$   
 $039035464356357669755894,$

$c_4 = 12040562557445113732473911652223993486907469290295457468128810943319836$   
 $89865239881378622819502882220194632147152122716053991772220199156765230$   
 $27510298197878920424163371390730684468025560664928791033866500650711867$   
 $64062267173825842140914417574375291482513842223759443700367430229129300$   
 $5181012290132879477619117,$

$c_5 = 66699227296798574343397095996822802356352501707241226477647274006905968$   
 $10673854955408065485554029221240872924799602606228672828557689341973222$   
 $51563295263105644219421197195443848039478669501756964029791476750411047$   
 $04900146183224578433038704943628899192782802758784232421703243761756962$   
 $894634288996400822908449,$

$c_6 = 35842810907626545862969792069041104823971684976414879250901086369792821$   
 $45580427757242397757516407715510081504624842581515104846891405331533253$   
 $61167679156090524981249642173885098604795315049563957046069182482249951$   
 $05623301480564801725462585729948675579745223831303862071442959713127023$   
 $481367052128284782369023,$

and

$c_7 = 47155671451031215409844124716042641165572494776610186942841230939281162$   
 $67241186266669860841735961194282501443823932882005051795195976808037886$   
 $40329386789302925870414294037418371443238801520123232628229674559133870$   
 $46744561955317811032439098967200824031961905675407869670714450791407371$   
 $569282383073692093702050.$

**Decrypt:** In this algorithm, we first merge these ciphertexts to evaluate  $A_0$  by Fiat's method.

$A_0 = 52359866444019503369219138689244129876435136131866887023807097362559424$   
 $85204585255108074606984294095293919326326201521808678582533044952634349$   
 $94510658185476933422534369085535424171842256268376777043427683909828453$   
 $16021327514525201378524377441114039247482047879358064064184921287633493$   
 $11895521745079888137070,$

After that, we take one decryption which does a full-scale modular exponentiation  $A \equiv A_0^d \pmod{N}$ :

$A = 97020381281501165266847566088201413180464241138685997123851463155123151$   
 $43976735868920221887471189785523640074792116012536463406101039184891861$   
 $98642917692578018622021094482005608364078487921673756050283179384743906$   
 $97318487470321735651171108807814620809108832763769829503607711597977756$   
 $914765803133937106665598.$

Finally, from Equation (4.24), we split the plaintext using the following  $\alpha_i$ .

$$\begin{aligned}\alpha_1 &= 1737436848943334, & \alpha_2 &= 1241026320673810, \\ \alpha_3 &= 789744022246970, & \alpha_4 &= 511010837924510, \\ \alpha_5 &= 55332383724310, & \alpha_6 &= 9039733865470, \\ \alpha_7 &= 1974979930.\end{aligned}$$



We get the plaintext from Equation (4.25).

$$\begin{aligned}
 m'_1 &\equiv \frac{c_1^{(\alpha_1+1)/e_1} \prod_{2 \leq j \leq 7} c_j^{\alpha_1/e_j}}{A^{\alpha_1}} \\
 &\equiv c_1^{347487369788667} \cdot c_2^{248205264134762} \cdot c_3^{157948804449394} \\
 &\quad \cdot c_4^{102202167584902} \cdot c_5^{11066476744862} \cdot c_6^{1807946773094} \\
 &\quad \cdot c_7^{394995986} \cdot A^{-1737436848943334} \pmod{N}
 \end{aligned}$$

We can find that  $m'_1 = m_1$ , which shows the correctness of our implementation. Note that  $m_2, m_3, \dots, m_7$  can be computed in the same way from Equation (4.25) and Equation (4.26). In fact, we compute  $m_2, m_3, \dots, m_7$  first and then compute  $m_1$  as Remark 4.4.1, we can finish the split (decryption) efficiently.

### 4.4.3 Implementation Based on Minus-Type Equations

We find that the solutions of the following Diophantine Equation

$$\sum_{1 \leq i \leq n} \frac{1}{x_i} - \prod_{1 \leq i \leq n} \frac{1}{x_i} = 1, \quad 2 \leq x_1 < x_2 < \dots < x_n \quad (4.33)$$

can also be used to implement batch RSA so that  $\alpha_i$ , which satisfies Equation (4.22), can be constructed directly and the whole decryption only uses modular  $N$ . We call Equation (4.33) the *Minus-Type Diophantine Equation* or *Minus-Type Equation*. For convenience, throughout this book, Equation (4.33) with  $n$  unknown variables is denoted as Equation (4.33) <sub>$n$</sub> .

Our new implementation of batch RSA also contains three algorithms: KeyGen, Encrypt, and Decrypt. The essence of the implementation is in the decryption algorithm. Algorithm KeyGen and Encrypt are the same as our implementation based on the Plus-Type Equations.

**Decrypt:** This algorithm takes a public key  $(N, b, e_1, e_2, \dots, e_b, x = \prod_{1 \leq i \leq n} x_i)$ , a private key  $d$ , and ciphertexts  $c_1, c_2, \dots, c_b$  as input. It evaluates  $A_0 \equiv \prod_{1 \leq i \leq b} c_i^{e/e_i}$

(mod  $N$ ) by the methods proposed by Fiat, and  $A \equiv A_0^d \pmod{N}$ . Obviously

$$A \equiv \prod_{1 \leq i \leq b} m_i \pmod{N}.$$

Then it evaluates

$$\alpha_i = \frac{x}{e_i}, \quad 1 \leq i \leq b,$$

and it outputs

$$m_i \equiv \frac{A^{\alpha_i}}{C_i^{(\alpha_i-1)/e_i} \cdot \prod_{j \neq i} C_j^{\alpha_i/e_j}} \pmod{N}, \quad 1 \leq i \leq b. \quad (4.34)$$

*Remark 4.4.4.* In our decrypt algorithm, after the plaintext  $(m_1, m_2, \dots, m_{b-1})$  are decrypted, the final  $m_b$  can be calculated by

$$m_b = \frac{A}{\prod_{1 \leq j < b} m_j} \pmod{N}$$

where computation can be reduced.

In our implementation,  $b - 1$  modulars are reduced, compared with Fiat's batch RSA and Boneh and Shacham's improvement. Since  $(x_1, x_2, \dots, x_n)$  is a solution of Equation (4.33) <sub>$n$</sub>  and  $\{e_1, e_2, \dots, e_b\} \subseteq \{x_1, x_2, \dots, x_n\}$ ,  $\alpha_1, \alpha_2, \dots, \alpha_b$  are evaluated by  $\alpha_i = \frac{\prod_{1 \leq j \leq n} x_j}{e_i}$  without modular inversion. Therefore, there is only one kind of modular operations, which is "mod  $N$ ." If we make a hardware circuit for batch RSA decryption, then our implementation requires one circuit module doing modular operations, and other two implementations both require  $b$  modules doing modular operations. Our implementation is more suitable for hardware circuit.

**Theorem 4.4.5** (Correctness). *The above scheme is correct.*

*Proof.* See the proof of Theorem 4.4.3. □

For a simple example, we choose a solution  $(2, 3, 11, 13)$  of Equation  $(4.33)_4$  and take 3, 11, and 13 as the public keys  $e_1$ ,  $e_2$ , and  $e_3$ , respectively.

First, we have

$$x = 2 \cdot 3 \cdot 11 \cdot 13 = 858,$$

and

$$e = e_1 \cdot e_2 \cdot e_3 = 429,$$

and

$$A \equiv (c_1^{429/3} \cdot c_2^{429/11} \cdot c_3^{429/13})^{1/429} \equiv (c_1^{143} \cdot c_2^{39} \cdot c_3^{33})^{1/429} \pmod{N}. \quad (4.35)$$

Second, we can get  $\alpha_i$ :

$$\alpha_1 = \frac{x}{e_1} = \frac{858}{3} = 286,$$

$$\alpha_2 = \frac{x}{e_2} = \frac{858}{11} = 78,$$

and

$$\alpha_3 = \frac{x}{e_3} = \frac{858}{13} = 66.$$

Third, we can get  $m_i$ :

$$m_1 \equiv \frac{A^{286}}{c_1^{95} \cdot c_2^{26} \cdot c_3^{22}} \pmod{N}, \quad (4.36)$$

$m_2$  and  $m_3$  can be obtained in the same way.

#### 4.4.4 A Concrete Example Based on Minus-Type Equations

We explain the implementation by a concrete example.

KeyGen: Let the batch size  $b = 4$ . We choose  $p = 107$  and  $q = 83$ . Let  $N = p \cdot q = 8881$ . From the solution  $\{x_1 = 2, x_2 = 3, x_3 = 11, x_4 = 17, x_5 = 59\}$  of

Equation (4.33))<sub>5</sub>, we choose the public key as follows:  $e_1 = 3, e_2 = 11, e_3 = 17$ , and  $e_4 = 59$ . Let  $e = \prod_{1 \leq i \leq b} e_i = 33099$  and  $x = \prod_{1 \leq i \leq 5} x_i = 66198$ . The private key is  $d \equiv e^{-1} \pmod{\phi(N)} \equiv 4083$ .

**Encrypt:** Suppose the plaintexts are  $m_1 = 1000, m_2 = 2000, m_3 = 3000$ , and  $m_4 = 4000$ . We compute the ciphertexts  $c_1 = 8281, c_2 = 1316, c_3 = 4824$ , and  $c_4 = 169$ .

**Decrypt:** The algorithm evaluates the following values. We have  $A_0 = 7441$  and take one decrypt operation  $A = A_0^{4083} \pmod{8881} = 4982$ . We can obtain  $\alpha_i$  as follows:

$$\begin{aligned} \alpha_1 &= \frac{66198}{3} = 22066, & \alpha_2 &= \frac{66198}{11} = 6018, \\ \alpha_3 &= \frac{66198}{17} = 3894, & \alpha_4 &= \frac{66198}{59} = 1122. \end{aligned}$$

Then the algorithm outputs:

$$m'_1 \equiv \frac{4982^{22066}}{8281^{(22066-1)/3} \times 1316^{22066/11} \times 4824^{22066/17} \times 169^{22066/59}} \pmod{8881} = 1000,$$

where it shows that  $m'_1 = m_1$ . Note that  $m'_2, m'_3$ , and  $m'_4$  can be computed in the same way:

$$\begin{aligned} m'_2 &\equiv \frac{4982^{6018}}{8281^{6018/3} \times 1316^{(6018-1)/11} \times 4824^{6018/17} \times 169^{6018/59}} \pmod{8881} = 2000 \\ m'_3 &\equiv \frac{4982^{3894}}{8281^{3894/3} \times 1316^{3894/11} \times 4824^{(3894-1)/17} \times 169^{3894/59}} \pmod{8881} = 3000 \\ m'_4 &\equiv \frac{4982^{1122}}{8281^{1122/3} \times 1316^{1122/11} \times 4824^{1122/17} \times 169^{(1122-1)/59}} \pmod{8881} = 4000. \end{aligned}$$

Obviously,  $m'_2 = m_2, m'_3 = m_3$ , and  $m'_4 = m_4$ .

## 4.5 Solving the Diophantine Equations

### 4.5.1 Plus-Type Equations

#### Algorithms for Solving the Plus-Type Equations

Equation (4.23) is a special case of the problem of expressing 1 as the sum of distinct unit fractions. Equation (4.23) also provides solutions to Zám's problem: find all sequences  $\{N_1, \dots, N_k\}$  of integers  $\geq 2$  with the property that for each  $i$ ,  $N_i$  properly divides  $1 + \prod_{1 \leq j \leq k, j \neq i} N_j$ . In the past forty years, many mathematical researchers have worked on this equation [50, 51, 53, 57, 58, 74–76, 109, 169, 263, 264, 267, 268, 270]. In recent years, applications of Equation (4.23) are also found in the graph theory [52, 58], differential geometry [49, 51], and computation theory [103].

However, it is impossible to solve the Plus-Type Equation when  $n \geq 8$  by hand. See Table 4.1 for the known possible value intervals for each variable ( $n = 8$ ).

Table 4.1: The known possible values

$x_i$	possible values
$x_1$	2
$x_2$	[ 3, 5 ]
$x_3$	[ 7, 13 ]
$x_4$	[ 11, 71 ]
$x_5$	[ 17, 3559 ]
$x_6$	[ 67, $3.5 \times 10^6$ ]
$x_7$	[ 551, $1.0 \times 10^{13}$ ]
$x_8$	[ $8.6 \times 10^4$ , $1.1 \times 10^{26}$ ]

Although the following results are known, Equation (4.23)<sub>8</sub> has not been solved completely until our work was published. Table 4.2 gives the number of solutions for  $n \leq 8$  in previous works. We have presented an algorithm to solve Equation (4.23) by elliptic curve method. We found all of the 122 solutions for  $n = 8$  in November, 2008. The computation was carried out on one personal computer with 2.3 GHz CPU

Table 4.2: The number of solutions for  $n \leq 8$  for Equation (4.23)<sub>n</sub>

Equations	Number of solutions	Year	Contributor
$n \leq 6$	15	1964	Ke and Sun [169]
$n = 7$	18	1978	Janak and Skula [158]
$n = 7$	26	1987	Cao, Liu and Zhang [75] [76]
$n = 7$	26	1988	Brenton and Hill [51]
$n = 8$	119	1998	Brenton and Vasilu [53]

and lasted 21 days. We have provided the algorithm and all 122 solutions. Although we have not solved Equation (4.23) for all the solutions when  $n = 9$ , we have generated solutions based on known solutions [70]. We have found 411 solutions for  $n = 9$ , and 2318 solutions for  $n = 10$ . Since our method is universal, it could be used to find a number of solutions for  $n > 10$ .

Now, we introduce the algorithm to solve Equation (4.23). Firstly, we state some lemmas about Equation (4.23). Then we present the outline and the pseudo codes of the algorithm. Finally we propose a new theorem.

**Lemma 4.5.1.** *Suppose that  $(x_1, x_2, \dots, x_n)$  is a solution of Equation (4.23)<sub>n</sub>. Then  $x_1, x_2, \dots, x_n$  are pairwise coprime.*

*Proof.* See [51]. □

**Lemma 4.5.2** (Sun and Cao [268]). *Let  $(x_1^{(j)}, x_2^{(j)}, \dots, x_{n-1}^{(j)})$  be the  $j$ -th solution of Equation (4.23)<sub>n-1</sub>,  $1 \leq j \leq \Omega(n-1)$ . Put*

$$k_j(n) = (x_1^{(j)} \cdot x_2^{(j)} \cdot \dots \cdot x_{n-1}^{(j)})^2 + 1, 1 \leq j \leq \Omega(n-1).$$

*Then*

$$\Omega(n+1) \geq \Omega(n) + \sum_{j=1}^{\Omega(n-1)} \left( \frac{d(k_j(n))}{2} - 1 \right)$$

*where  $d(k_j(n))$  denotes the number of different positive factors of  $k_j(n)$ .*

*Proof.* The proof consists of two steps:

**Step 1:** Let  $(x_1, x_2, \dots, x_n)$  be any solution of Equation (4.23)<sub>n</sub>. Put  $x_{n+1} = \prod_{1 \leq i \leq n} x_i + 1$ . By a direct computation,  $(x_1, x_2, \dots, x_{n+1})$  is a solution of Equation (4.23)<sub>n+1</sub>. Hence,  $\Omega(n+1) \geq \Omega(n)$ .

**Step 2:** Let  $(x_1^{(j)}, x_2^{(j)}, \dots, x_{n-1}^{(j)})$  be as above. Let  $k$  be one positive factor of  $k_j(n)$ , such that  $1 < k < \sqrt{k_j(n)}$ . Put

$$x_n^{(j)} = \prod_{1 \leq i \leq n-1} x_i + k, x_{n+1}^{(j)} = \prod_{1 \leq i \leq n-1} x_i + \frac{k_j(n)}{k}.$$

By a direct computation, it is proved that  $(x_1^{(j)}, x_2^{(j)}, \dots, x_{n+1}^{(j)})$  is a solution of Equation (4.23)<sub>n+1</sub>. For each  $j$ ,  $1 \leq j \leq \Omega(n-1)$ , the number of  $k$  satisfying  $1 < k < \sqrt{k_j(n)}$  is  $\frac{d(k_j(n))}{2} - 1$ . Hence, we can generate  $\sum_{j=1}^{\Omega(n-1)} \left( \frac{d(k_j(n))}{2} - 1 \right)$  solutions of Equation (4.23)<sub>n+1</sub>. These solutions are all different from the solutions generated in Step 1, since  $x_n^{(j)} \neq \prod_{1 \leq i \leq n-1} x_i + 1$ . Therefore,  $\Omega(n+1) \geq \Omega(n) + \sum_{j=1}^{\Omega(n-1)} \left( \frac{d(k_j(n))}{2} - 1 \right)$ .  $\square$

**Lemma 4.5.3.** Suppose  $2 \leq x_1 < \dots < x_{n-2}$  and  $x_1, x_2, \dots, x_{n-2}$  are pairwise coprime. Put

$$A = \prod_{1 \leq i \leq n-2} x_i, B = A - \sum_{1 \leq i \leq n-2} \frac{A}{x_i}.$$

Let  $P$  and  $Q$  be positive integers such that  $PQ = A^2 + B$ ,  $P < Q$ . Put

$$x_{n-1} = \frac{A+P}{B}, x_n = \frac{A+Q}{B}.$$

If  $B|A+P$  and  $B|A+Q$ ,  $(x_1, x_2, \dots, x_n)$  is a solution to Equation (4.23)<sub>n</sub>.

*Proof.* This lemma is proved by a direct computation.  $\square$

**Lemma 4.5.4.** Put  $x_1 = \mathcal{A}$ , where  $\mathcal{A}$  is an arbitrary positive integer. For  $1 \leq i < n$ , let  $x_{i+1}$  be the minimum positive integer such that  $x_{i+1} > x_i$  and  $x_{i+1}$  is coprime to

all  $x_1, x_2, \dots, x_i$ . If

$$\frac{1}{\mathcal{A}} + \sum_{1 \leq i \leq n} \frac{1}{x_i} + \frac{1}{\mathcal{A}} \prod_{1 \leq i \leq n} \frac{1}{x_i} < 1,$$

there is no solution of Equation (4.23)<sub>n</sub> with  $x_1 = \mathcal{A}$ .

*Proof.* Suppose there is one solution of Equation (4.23)<sub>n</sub> with  $x_1 = \mathcal{A}$ . For  $1 \leq i < n$ ,  $x_{i+1}$  is the minimum positive integer such that  $x_{i+1} > x_i$  and  $x_{i+1}$  is coprime to  $x_1, x_2, \dots, x_i$ .

For all solutions  $(\mathcal{A}, x_2^*, x_3^*, \dots, x_n^*)$  of Equation (4.23)<sub>n</sub>, we have

$$\sum_{1 \leq i \leq n} \frac{1}{x_i} + \prod_{1 \leq i \leq n} \frac{1}{x_i} \geq \sum_{1 \leq i \leq n} \frac{1}{x_i^*} + \prod_{1 \leq i \leq n} \frac{1}{x_i^*} = 1$$

It is a contradiction. □

### Algorithm Outline—Direct Finding Algorithms

Let  $(x_1, x_2, \dots, x_n)$  be a solution of Equation (4.23)<sub>n</sub>.

- First we calculate the lower bound of  $x_1$ ,  $L_1$ , and the upper bound of  $x_1$ ,  $U_1$ .

Let  $x_1$  be  $\mathcal{A}$ , and  $x_{i+1}$  be the minimum positive integer such that  $x_{i+1} > x_i$ , and  $x_{i+1}$  is coprime to  $x_1, x_2, \dots, x_i$ ,  $1 \leq i < n$ . If  $\sum_{1 \leq i \leq n} \frac{1}{x_i} + \prod_{1 \leq i \leq n} \frac{1}{x_i} \geq 1$ , then  $\mathcal{A} \in [L_1, U_1]$  by Lemma 4.5.17.

For example, for Equation (4.23)<sub>8</sub>, we fix  $x_1 = 2$  and then we get  $x_2 = 3$ ,  $x_3 = 5$ ,  $x_4 = 7$ ,  $x_5 = 11$ ,  $x_6 = 13$ ,  $x_7 = 17$ , and  $x_8 = 19$ . From

$$\frac{1}{2} + \frac{1}{3} + \frac{1}{5} + \frac{1}{7} + \frac{1}{11} + \frac{1}{13} + \frac{1}{17} + \frac{1}{19} + \frac{1}{2 \cdot 3 \cdot 5 \cdot 7 \cdot 11 \cdot 13 \cdot 17 \cdot 19} > 1$$

we have  $2 \in [L_1, U_1]$ .

- Then for all values of  $x_1$  over the range  $[L_1, U_1]$ , we calculate the lower bound of



$x_2$ ,  $L_2$ , and the upper bound of  $x_2$ ,  $U_2$ . In this way we fix  $(x_1, x_2, \dots, x_i)$ , then we calculate  $L_{i+1}$  and  $U_{i+1}$ ,  $1 \leq i < n-1$ .

– From

$$\frac{1}{x_1} + \dots + \frac{1}{x_i} < \sum_{1 \leq j \leq n} \frac{1}{x_j} + \prod_{1 \leq j \leq n} \frac{1}{x_j} = 1,$$

we have

$$x_i > \frac{\prod_{1 \leq k \leq i-1} x_k}{\prod_{1 \leq k \leq i-1} x_k - \sum_{1 \leq k \leq i-1} \prod_{1 \leq j \leq i-1, j \neq k} x_j}.$$

Hence,

$$L_i = \max \left\{ x_{i-1} + 1, \left\lceil \frac{\prod_{1 \leq k \leq i-1} x_k}{\prod_{1 \leq k \leq i-1} x_k - \sum_{1 \leq k \leq i-1} \prod_{1 \leq j \leq i-1, j \neq k} x_j} \right\rceil \right\}.$$

– It follows from

$$\sum_{1 \leq k < i} \frac{1}{x_k} + (n-i+1) \frac{1}{x_i} > \sum_{1 \leq i \leq n} \frac{1}{x_i} + \prod_{1 \leq i \leq n} \frac{1}{x_i} = 1$$

that

$$x_i < \frac{\left( \prod_{1 \leq k \leq i-1} x_k \right) (n-i+1)}{\prod_{1 \leq k \leq i-1} x_k - \sum_{1 \leq k \leq i-1} \prod_{1 \leq j \leq i-1, j \neq k} x_j}.$$

Therefore,

$$U_i = \left\lfloor \frac{\left( \prod_{1 \leq k \leq i-1} x_k \right) (n-i+1)}{\prod_{1 \leq k \leq i-1} x_k - \sum_{1 \leq k \leq i-1} \prod_{1 \leq j \leq i-1, j \neq k} x_j} \right\rfloor.$$

– When  $(x_1, x_2, \dots, x_{i-1})$  are fixed, put  $x_i = U_i$  and let  $x_j$  ( $i < j \leq n$ ) be the minimum positive integer such that  $x_j$  is coprime to  $x_1, \dots, x_{j-1}$  and  $x_j > x_{j-1}$ . Then  $\sum_{1 \leq i \leq n} \frac{1}{x_i} + \prod_{1 \leq i \leq n} \frac{1}{x_i} \geq 1$ . We can further reduce  $U_i$  based on this fact.

- After all the lower/upper bounds of  $(x_1, x_2, \dots, x_{n-2})$  are calculated, for all values of  $x_i$  over the range  $[L_i, U_i]$  ( $1 \leq i \leq n-2$ ), put  $A = \prod_{1 \leq i \leq n-2} x_i$ ,  $B = A - \sum_{1 \leq i \leq n-2} \frac{A}{x_i}$ . Factorize  $A^2 + B$ . For all positive integers  $P$  and  $Q$  such that  $PQ = A^2 + B$ ,  $P < Q$ , put  $x_{n-1} = \frac{A+P}{B}$ ,  $x_n = \frac{A+Q}{B}$ . If both  $x_{n-1}$  and  $x_n$  are integers,  $(x_1, x_2, \dots, x_n)$  is a solution of Equation (4.23)<sub>n</sub>.

There are many methods to factorize  $A^2 + B$ , such as general number field sieve, multiple polynomial quadratic sieve, elliptic curve method, and quantum algorithm. The general number field sieve [56, 88] is the most efficient algorithm known for factoring integers larger than 100 digits. Its complexity is  $L_N[\frac{1}{3}, c]$ , where  $L_N[c_1, c_2] = \mathcal{O}(e^{(c_2+o(1))(\log N)^{c_1}(\log \log N)^{1-c_1}})$ , and  $N$  is the integer to be factorized. The complexity of multiple polynomial quadratic sieve is  $L_N[\frac{1}{2}, 1]$ . The elliptic curve method [103] is the most suitable method for finding small factors. Its running time is dominated by the size of the smallest factor, while the complexity of the first two methods is both determined by the size of the number to be factored.

In our algorithm, integers are factorized by the elliptic curve method. Other methods also work well.

## Pseudo Codes

We present the pseudo codes of algorithm for further explanation.

---

### Algorithm 1 Search1

---

#### Procedure Search1(*depth*)

**Input:** *depth*

**Output:** All solutions of Equation (4.23)<sub>n</sub>.

1: **if** *depth* = *n* **then**

2: 
$$x_n \leftarrow \frac{\prod_{1 \leq k \leq n-1} x_k + 1}{\prod_{1 \leq k \leq n-1} x_k - \sum_{1 \leq k \leq n-1} \prod_{1 \leq i \leq n-1, i \neq k} x_i}$$

```

3:      if  $x_n$  is an integer then
4:          output  $x_1, x_2, \dots, x_n$ 
5:      end if
6:  end if
7:  if  $depth \leq n$  then
8:       $L_{depth} \leftarrow \max \left\{ x_{depth-1} + 1, \right.$ 
           $\left. \left\lceil \frac{\prod_{1 \leq k \leq depth-1} x_k}{\prod_{1 \leq k \leq depth-1} x_k - \sum_{1 \leq k \leq depth-1} \prod_{1 \leq j \leq depth-1, j \neq k} x_j} \right\rceil \right\}$ 
9:       $U_{depth} \leftarrow \lfloor \frac{(\prod_{1 \leq k \leq depth-1} x_k)(n - depth + 1)}{\prod_{1 \leq k \leq depth-1} x_k - \sum_{1 \leq k \leq depth-1} \prod_{1 \leq j \leq depth-1, j \neq k} x_j} \rfloor$ 
10:     call ReduceU( $depth$ )
11:     for all  $x_{depth} \in [L_{depth}, U_{depth}]$  do
12:         if  $(x_1, \dots, x_{depth})$  are all relatively prime then
13:             Call Search1( $depth + 1$ )
14:         end if
15:     end for
16: end if
17: if  $depth = n - 1$  then
18:      $L_{depth} \leftarrow \max \left\{ x_{depth-1} + 1, \right.$ 
           $\left. \left\lceil \frac{\prod_{1 \leq k \leq depth-1} x_k}{\prod_{1 \leq k \leq depth-1} x_k - \sum_{1 \leq k \leq depth-1} \prod_{1 \leq j \leq depth-1, j \neq k} x_j} \right\rceil \right\}$ 
19:      $U_{depth} \leftarrow \lfloor \frac{(\prod_{1 \leq k \leq depth-1} x_k)(n - depth + 1)}{\prod_{1 \leq k \leq depth-1} x_k - \sum_{1 \leq k \leq depth-1} \prod_{1 \leq j \leq depth-1, j \neq k} x_j} \rfloor$ 
20:     call ReduceU( $depth$ )
21:     if  $U_{depth} - L_{depth} < 120000$  then
22:         for all  $x_{depth} \in [L_{depth}, U_{depth}]$  do
23:             if  $(x_1, \dots, x_{depth})$  are all relatively prime then
24:                 call Search1( $depth + 1$ )
25:             end if
26:         end for
27:     else

```

```

28:       $A \leftarrow \prod_{1 \leq i \leq n-2} x_i$ 
29:       $B \leftarrow A - \sum_{1 \leq i \leq n-2} \frac{A}{x_i}$ 
30:       $C \leftarrow A^2 + B$ 
31:      for all positive factors  $P$  of  $C$  do
32:           $Q \leftarrow \frac{C}{P}$ 
33:          if  $P < Q$  then
34:               $x_{n-1} \leftarrow \frac{A+P}{B}$ 
35:               $x_n \leftarrow \frac{A+Q}{B}$ 
36:              if  $x_{n-1} > x_{n-2}$  and both  $x_{n-1}$  and  $x_n$  are integers then
37:                  output  $x_1, x_2, \dots, x_n$ 
38:              end if
39:          end if
40:      end for
41:  end if
42:  end if

```

---



---

**Function 1**  $reduceU(depth)$ 


---

**Input:**  $depth$

**Output:** Nothing.

**Step 1.**  $x_{depth} \leftarrow U_{depth}$

**Step 2.**  $x_{depth+1} \leftarrow$  the minimum value such that  $(x_1, \dots, x_{depth+1})$  are all relatively prime.

$x_{depth+2} \leftarrow$  the minimum value such that  $(x_1, \dots, x_{depth+2})$  are all relatively prime.

...

$x_n \leftarrow$  the minimum value such that  $(x_1, \dots, x_n)$  are all relatively prime.

**Step 3.** If  $\sum_{1 \leq i \leq n} \frac{1}{x_i} + \prod_{1 \leq i \leq n} \frac{1}{x_i} < 1$ , then

$$U_{depth} \leftarrow U_{depth} - 1$$

Goto Step 1.

End if

---

*Remark 4.5.5.* The complete solutions of Equation (4.23) can be found by running Algorithm Search1( $depth = 1$ ).

*Remark 4.5.6.* Algorithm *search* can be further optimized. The computation amount of decomposing an integer is about 60,000 times as the computation amount of computing  $x_n$  based on  $(x_1, \dots, x_{n-1})$ . Hence, we factorize  $A^2 + B$  to get  $x_{n-1}, x_n$  only if  $U_{n-1} - L_{n-1} > 120,000$ .

*Remark 4.5.7.* We apply the computer program in [4] to factorization. We would like to thank D. Alpern again.

The computer program examines all possible tuple  $(x_1, \dots, x_6)$ , and executes the factorization to find  $x_7$  and  $x_8$  for each tuple. The maximum value to be factorized is about  $10^{26}$ , which can be factorized efficiently by elliptic curve method. Elliptic curve method is a randomized method. It may fail to factorize by treating a composite number as a prime. We make sure that this method is always successful when the computer continues searching for the solutions.

**Theorem 4.5.8.** Equation (4.23)<sub>8</sub> has only 122 solutions. There is only one solution such that  $x_1, \dots, x_8$  are all primes: 2, 3, 11, 23, 31, 47059, 2217342227, 1729101023519.

All these 122 solutions can be found in Appendix A.

### New Solutions Generated from Known Solutions

We propose a new method, which can generate solutions of Equation (4.23) <sub>$n+3$</sub>  from solutions of Equation (4.23) <sub>$n$</sub>  [68–70]. We call it *Gap-three extension* algorithm. See the following pseudo codes.

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**Algorithm 2 Gap-three Extension**


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**Procedure** GapThreeExtension

**Input:**  $(x_1, x_2, \dots, x_n)$ , one solution of Equation (4.23) <sub>$n$</sub> ;

$LOW$ , lower bound of  $t$ ;  $HIGH$ , upper bound of  $t$ .

**Output:** Solutions of Equation (4.23) <sub>$n+3$</sub>

```

1:   $J \leftarrow \prod_{1 \leq i \leq n} x_i$ 
2:   $t \leftarrow 1$ 
3:  while  $t < LOW$  do
4:       $t \leftarrow t + 4$ 
5:  end while
6:  while  $t + 4 \leq HIGH$  do
7:      if  $t$  is not coprime to any of  $x_1, x_2, \dots, x_n$  then
8:          Continue
9:      end if
10:     Factorize  $J^2 + t$ 
11:     for each odd positive factor  $k$  of  $J^2 + t$  do
12:         if  $k \not\equiv 1 \pmod{4}$  or  $k = 1$  or  $k$  is not coprime with any of
             $(x_1, x_2, \dots, x_n)$  then
13:             continue
14:         end if
15:          $x_{n+1} \leftarrow J + k$ 
16:          $x_{n+2} \leftarrow J + \frac{J^2 + t}{k}$ 
17:         if  $x_{n+1} \geq x_{n+2}$ , or  $x_{n+1}$  and  $x_{n+2}$  are not coprime then
18:             continue
19:         end if
20:         if  $t \nmid [(J^2 + J \cdot k)^2 + k]$  then
```

```

21:          continue
22:      end if
23:       $x_{n+3} \leftarrow \frac{J \cdot x_{n+1} \cdot x_{n+2} + 1}{t}$ 
24:      if  $x_{n+2} \geq x_{n+3}$ , or  $x_{n+3}$  is not coprime with any of
           $(x_1, x_2, \dots, x_{n+2})$  then
25:          continue
26:      end if
27:      return  $(x_1, x_2, \dots, x_{n+3})$ 
28:  end for
29:   $t \leftarrow t + 4$ 
30: end while

```

---

The correctness of this method is based on the following lemma.

**Lemma 4.5.9** (Cao [70]). *Let  $(x_1, x_2, \dots, x_{n-2})$  be a solution of Equation (4.23) $_{n-2}$ . Put  $N = \prod_{1 \leq i \leq n-2} x_i$ . Then*

$$\begin{aligned}
 (x_1, x_2, \dots, x_{n-2}, x_{n-1} = N + k, x_n &= N + \frac{N^2 + t}{k}, x_{n+1} \\
 &= \frac{1}{t} [N(N + k)(N + \frac{N^2 + t}{k}) + 1])
 \end{aligned}$$

*is a solution of Equation (4.23) $_{n+1}$ , where  $x_{n-1}, x_n, x_{n+1}$  are integers satisfying  $x_{n-1} < x_n < x_{n+1}$  and  $t, k$  are positive integers. Moreover,  $N, k, t$  are pairwise coprime. If  $2|N, k \equiv t \equiv 1 \pmod{4}$ .*

*Proof.* The proof consists of the following steps:

**Step 1:** By a direct computation, it is easy to check that  $(x_1, x_2, \dots, x_{n+1})$  is a solution of Equation (4.23) $_{n+1}$ .

**Step 2:** Set  $y = x_{n-1}, z = x_n, w = x_{n+1}$ . Then from

$$1 = \sum_{1 \leq i \leq n+1} \frac{1}{x_i} + \prod_{1 \leq i \leq n+1} \frac{1}{x_i},$$

we have

$$1 = 1 - \frac{1}{N} + \frac{1}{y} + \frac{1}{z} + \frac{1}{w} + \frac{1}{N \cdot y \cdot z \cdot w}.$$

Therefore,

$$\frac{k}{N \cdot (N+k)} = \frac{1}{z} + \frac{1}{w} + \frac{1}{N \cdot (N+k) \cdot z \cdot w}.$$

Hence,

$$z \cdot w \cdot k = N \cdot (N+k) \cdot w + N \cdot (N+k) \cdot z + 1.$$

Let  $d$  denote  $\gcd(N, k)$ . Then  $z \cdot w \cdot k \equiv 0 \pmod{d}$ , and  $N \cdot (N+k) \cdot w + N \cdot (N+k) \cdot z + 1 \equiv 1 \pmod{d}$ . Therefore,  $d|1$ . Hence,  $N$  and  $k$  are coprime.

**Step 3:** Let  $d'$  denote  $\gcd(k, t)$ . From  $w \in \mathbb{Z}$ , we have  $t|[N \cdot (N+k)(N + \frac{N^2+t}{k}) + 1]$ . Hence,  $d'|[N \cdot (N+k)(N + \frac{N^2+t}{k}) + 1]$ . From  $d'|k$ , we get  $d'|[N^2(N + \frac{N^2+t}{k}) + 1]$ . Since  $\frac{N^2+t}{k} \in \mathbb{Z}$ , we have  $k|(N^2 + t)$ . Therefore  $d'|N^2$ . It follows that  $d'|1$ . Hence,  $k$  and  $t$  are coprime.

**Step 4:** Let  $d''$  denote  $\gcd(N, t)$ . From

$$w = \frac{1}{t} \left[ N \cdot (N+k) \left( N + \frac{N^2+t}{k} \right) + 1 \right],$$

it follows that

$$w \cdot t = N \cdot (N+k) \left( N + \frac{N^2+t}{k} \right) + 1.$$

Since  $d''|(wt)$  and  $d''|N$ , we have  $d''|1$ . Therefore,  $N$  and  $t$  are coprime.

It follows from Steps 2, 3, and 4 that  $N$ ,  $k$ , and  $t$  are pairwise coprime.

**Step 5:** Assume that  $2|N$ . Since  $t|[N \cdot (N+k) \cdot (N + \frac{N^2+t}{k}) + 1]$  and  $2|N$ , we have  $2 \nmid t$ . From  $k|(N^2 + t)$ , we have  $2 \nmid k$ . From  $(k, t) = 1$ , we get  $(\frac{-t}{k}) = 1$ . Then from  $t|[N \cdot (N+k) \cdot (N + \frac{N^2+t}{k}) + 1]$ , we have  $t|[N \cdot (N+k) \cdot (nk + N^2 + t) + k]$ , i.e.,  $t|[N^2 \cdot (N+k)^2 + k]$ . Therefore  $(\frac{-k}{t}) = 1$ . We have,

$$1 = \left( \frac{-t}{k} \right) \left( \frac{-k}{t} \right) = (-1)^{\frac{k-1}{2} + \frac{t-1}{2}} \left( \frac{t}{k} \right) \left( \frac{k}{t} \right) = (-1)^{\frac{k-1}{2} + \frac{t-1}{2} + \frac{k-1}{2} \cdot \frac{t-1}{2}}.$$



Thus,

$$\frac{k-1}{2} + \frac{t-1}{2} + \frac{k-1}{2} \cdot \frac{t-1}{2} \equiv 0 \pmod{2}.$$

$$k \equiv t \equiv 1 \pmod{4}.$$

□

Now we propose the following new theorem.

**Theorem 4.5.10.** *Let  $\Omega(k)$  denote the number of solutions of Equation (4.23)<sub>k</sub>. We have*

$$\Omega(n+3) \geq \Omega(n+2) + A(n+1) + B(n) + \Gamma(n+2),$$

where  $A(n+1) = \sum_{j=1}^{\Omega(n+1)} \left( \frac{d(k_j(n+2))}{2} - 1 \right)$ ,  $d(k_j(n+2))$  is defined in Lemma 4.5.2,  $B(n)$  denotes the number of solutions found by the Algorithm 2 based on  $n$ , restricted by  $t > 1$  and  $k > 1$ ,  $\Gamma(n+2)$  denotes the number of solutions generated as follows:  $(x_1, x_2, \dots, x_{n+2})$  be a solution of Equation (4.23)<sub>n+2</sub> satisfying that  $(x_1, x_2, \dots, x_{n+1})$  is not a solution of Equation (4.23)<sub>n+1</sub>, and  $C = (\prod_{1 \leq i \leq n+1} x_i)^2 + 1$ . Let  $k$  be a factor of  $C$ , such that  $1 < k < \sqrt{C}$ . Put

$$x'_{n+2} = \prod_{1 \leq i \leq n+1} x_i + k, x'_{n+3} = \prod_{1 \leq i \leq n+1} x_i + \frac{C}{k}.$$

Then

$$(x_1, x_2, \dots, x_{n+1}, x'_{n+2}, x'_{n+3})$$

is a solution of Equation (4.23)<sub>n+3</sub>.

*Proof.* Let  $S(k)$  be the set of all solutions of Equation (4.23)<sub>k</sub>. In fact there are at least four different methods to generate solutions of Equation (4.23)<sub>n+3</sub> listed in Table 4.6. Methods 3 and 4 are from [70]. They are used firstly to generate solutions of Equation (4.23) when  $n = 9, 10$  in Theorems 4.5.11 and 4.5.12.

Table 4.6: Four methods to generate solutions of Equation (4.23)<sub>n+3</sub>

Methods	Number of solutions	Brief illustration
Method 1: Gap-one extension	$\Omega(n+2)$	$x_1, x_2, \dots, x_{n+2},$ $x_{n+3} = (\prod_{1 \leq i \leq n+2} x_i) + 1.$ $(x_1, x_2, \dots, x_{n+2}) \in S(n+2).$
Method 2: Gap-two extension	$A(n+1)$	$x_1, x_2, \dots, x_{n+1}, x_{n+2} = J + k, x_{n+3} = J + \frac{J^2+1}{k}.$ $(x_1, x_2, \dots, x_{n+1}) \in S(n+1).$ $J = \prod_{1 \leq i \leq n+1} x_i, k   (J^2 + 1), 1 < k < \sqrt{J^2 + 1}.$
Method 3: Gap-three extension	$B(n)$	$x_1, x_2, \dots, x_n, x_{n+1} = N + k, x_{n+2} = N + \frac{N^2+t}{k},$ $x_{n+3} = \frac{1}{t} (\prod_{1 \leq i \leq n+2} x_i + 1), (x_1, x_2, \dots, x_n) \in S(n).$ $N = \prod_{1 \leq i \leq n} x_i, k > 1, t > 1,$ $x_{n+1} \in \mathbb{Z}, x_{n+2} \in \mathbb{Z}, x_{n+3} \in \mathbb{Z}.$
Method 4: (Another gap-one extension)	$\Gamma(n+2)$	$x_1, x_2, \dots, x_{n+1}, x'_{n+2} = J + k, x'_{n+3} = J + \frac{J^2+1}{k}.$ $(x_1, x_2, \dots, x_{n+2}) \in S(n+2),$ $(x_1, x_2, \dots, x_{n+1}) \notin S(n+1),$ $J = \prod_{1 \leq i \leq n+1} x_i, k   (J^2 + 1), 1 < k < \sqrt{J^2 + 1}.$

We prove that the solutions generated by the four methods are all different. The proof consists of the following three steps.

**Step 1:** It follows from Lemma 4.5.2 that the solutions generated by Method 2 are different from Method 1.

**Step 2:** The solutions generated by Method 3 are all different from Method 1. Because  $t > 1$ ,  $x_{n+3} < \prod_{1 \leq i \leq n+2} x_i + 1$  in Method 3, while  $x_{n+3} = \prod_{1 \leq i \leq n+2} x_i + 1$  in Method 1.

The solutions generated by Method 3 are also different from Method 2. Since  $k > 1$ ,  $(x_1, x_2, \dots, x_{n+1}) \notin S(n+1)$  in Method 3, while  $(x_1, x_2, \dots, x_{n+1}) \in S(n+1)$  in Method 2.

**Step 3:** The solutions generated by Method 4 are all different from Methods 2 and 3, because  $(x_1, x_2, \dots, x_{n+1}) \notin S(n+1)$ ,  $(x_1, x_2, \dots, x_n) \notin S(n)$  in Method 4, but  $(x_1, x_2, \dots, x_{n+1}) \in S(n+1)$  in Method 2,  $(x_1, x_2, \dots, x_n) \in S(n)$  in Method 3.

The solutions generated by Method 4 are also different from Method 1. Since  $k > 1$ ,  $x_{n+3} < \prod_{1 \leq i \leq n+2} x_i + 1$  in Method 4, while  $x_{n+3} = \prod_{1 \leq i \leq n+2} x_i + 1$  in Method 1.

Therefore, we have  $\Omega(n+3) \geq \Omega(n+2) + A(n+1) + B(n) + \Gamma(n+2)$ .  $\square$

**The number of solutions for  $n = 9, 10$**

**Theorem 4.5.11.** *There are at least 411 solutions of Equation (4.23)<sub>9</sub>.*

We found 411 solutions of Equation (4.23)<sub>9</sub>. The number of solutions found by each method is listed in the following Table 4.7

Table 4.7: Solutions found in Equation (4.23)<sub>9</sub>.

Solutions found at least	Method
122	Method 1, $\Omega(8)$
205	Method 2, $A(7)$
47	Method 3, $B(6)$
14	Method 4, $\Gamma(8)$
23	Direct finding

Because of space limitations, these solutions are listed in

<http://tdt.sjtu.edu.cn/9plus.html> and solutions found by different methods are also listed separately there.

**Theorem 4.5.12.** *There are at least 2318 solutions of Equation (4.23)<sub>10</sub>.*

We also found 2318 solutions of Equation (4.23)<sub>10</sub>. The number of solutions found by each method is listed in Table 4.8

Table 4.8: Solutions found in Equation (4.23)<sub>10</sub>.

Solutions found at least	Method
411	Method 1, $\Omega(9)$
1623	Method 2, $A(8)$
284	Method 3, $B(7)$

These solutions are listed in <http://tdt.sjtu.edu.cn/10plus.html>.

Solutions found by different methods are also listed separately there.

### 4.5.2 Minus-Type Equations

Similar to the Plus-Type Equations, we propose a fast algorithm to solve Equation (4.33) in this section.

#### Algorithms for Solving the Minus-Type Equations

Traditional algorithms to solve Equation (4.33) have high computational complexity. Our algorithm is based on factorizing integers using the elliptic curve method. The elliptic curve method is a fast, sub-exponential running time algorithm. We found all 550 solutions for  $n = 8$ , and we list them in this section. But the computation amount for  $n = 9$  is beyond our computational capacity. We cannot solve Equation (4.33) for  $n = 9$  directly. Therefore, we propose new methods to generate solutions from known solutions of Equation (4.23) with less variables. We have found 1547 solutions for  $n = 9$  and 18984 solutions for  $n = 10$ . Due to space limitations, these solutions are published at webpage <http://tdt.sjtu.edu.cn/9minus.html> and <http://tdt.sjtu.edu.cn/10minus.html>. These solutions can be used in the new batch RSA implementation.

Table 4.9: Research results about Equation (4.33)

Results	Researchers	Year
$n \geq 9 \rightarrow A(n+1) \geq \Omega(n) + \Omega(n-1) + 6$	Sun and Cao [266]	1985
$2 \nmid n \geq 9 \rightarrow A(n+1) \geq \Omega(n) + \Omega(n-1) + 10$	Sun and Cao [266]	1985
found the complete solutions for $n \leq 6$	Sun and Cao [265]	1986
$n \geq 9 \rightarrow A(n+1) \geq \Omega(n) + \Omega(n-1) + 10$	Sun and Cao [267]	1986
$2 \mid n \geq 12 \rightarrow A(n+1) \geq \Omega(n) + \Omega(n-1) + 14$	Sun and Cao [267]	1986
$n \geq 10 \rightarrow A(n+1) \geq \Omega(n) + \Omega(n-1) + 16$	Cao, Liu and Zhang [75]	1987
$2 \mid n \geq 12 \rightarrow A(n+1) \geq \Omega(n) + \Omega(n-1) + 18$	Cao, Liu and Zhang [75]	1987
$n \geq 10 \rightarrow A(n+1) \geq \Omega(n) + \Omega(n-1) + 34$	Cao [69]	1988
$2 \mid n \geq 12 \rightarrow A(n+1) \geq \Omega(n) + \Omega(n-1) + 46$	Cao [69]	1988

Equation (4.33) also helps to find Giuga numbers. A number  $K$  is a Giuga number

if and only if  $p | (\frac{K}{p} - 1)$  for all prime divisors  $p$  of  $K$ , or  $\sum_{p|K} \frac{1}{p} - \prod_{p|K} \frac{1}{p} \in \mathbb{N}$  [124]. Giuga numbers are named after Giuga [124], who formulated a conjecture on primality in 1950:  $N$  is prime if and only if  $\sum_{k=1}^{N-1} k^{N-1} \equiv 1 \pmod{N}$ . In 1996, Borwein et al. [44] published 11 Giuga numbers. Then Hogan and Mangilin found the 12th Giuga number [58], and Girsensohn found another one [58] in 2000. Sloane presented 13 Giuga numbers [261]. We prove that there are only 12 Giuga numbers with less than eight distinct prime factors, through the complete solutions of Equation (4.33) for  $n \leq 8$ , which are listed in Appendix B and Appendix C. The basic idea is that the equation

$$\sum_{p|K} \frac{1}{p} - \prod_{p|K} \frac{1}{p} = 1 \quad (4.37)$$

is a special case of Equation (4.33).  $\prod_{1 \leq i \leq n} x_i$  is a Giuga number if and only if  $(x_1, x_2, \dots, x_n)$  is a solution of Equation (4.33)<sub>n</sub> and  $x_1, x_2, \dots, x_n$  are all primes.

In the following section, we propose a fast algorithm to solve Equation (4.33) and a method to generate new solutions of Equation (4.33) from known solutions of Equation (4.23). We first state some useful lemmas.

**Lemma 4.5.13.** *If  $(x_1, x_2, \dots, x_n)$  is a solution of Equation (4.33), then  $x_1, x_2, \dots, x_n$  are pairwise coprime.*

*Proof.* For any  $i, j, 1 \leq i < j \leq n$ , put  $d = \gcd(x_i, x_j)$ . We have

$$\prod_{1 \leq i \leq n} x_i \left[ \sum_{1 \leq i \leq n} \frac{1}{x_i} - \prod_{1 \leq i \leq n} \frac{1}{x_i} \right] = \prod_{1 \leq i \leq n} x_i, d \prod_{1 \leq i \leq n} x_i.$$

Therefore,

$$d \prod_{1 \leq i \leq n} x_i \left[ \sum_{1 \leq i \leq n} \frac{1}{x_i} - \prod_{1 \leq i \leq n} \frac{1}{x_i} \right].$$

Hence,  $d|1$ , and so  $d = 1$ . This completes the proof.  $\square$

**Lemma 4.5.14.** *If  $(x_1, x_2, \dots, x_n)$  is a solution of Equation (4.33), then*

$$\frac{\prod_{1 \leq j \leq n} x_j}{x_i} \equiv 1 \pmod{x_i}, 1 \leq i \leq n.$$

*Proof.* By a direct computation, it is easy to prove the lemma.  $\square$

**Lemma 4.5.15** (Sun and Cao [266]). *Let  $(x_1^{(j)}, x_2^{(j)}, \dots, x_{n-1}^{(j)})$  be the  $j$ -th solution of Equation (4.23) $_{n-1}$ ,  $1 \leq j \leq \Omega(n-1)$ . Put  $l_j(n) = (\prod_{1 \leq i \leq n-1} x_i^{(j)})^2 - 1$ . We have*

$$A(n+1) \geq \Omega(n) + \sum_{1 \leq j \leq \Omega(n-1)} \left( \frac{d(l_j(n))}{2} - 1 \right),$$

where  $d(l_j(n))$  denote the number of different positive factors of  $l_j(n)$ .

*Proof.* The proof consists of two steps (Cao [70]) :

**Step 1:** Let  $k$  be a positive factor of  $l_j(n)$  satisfying  $1 < k < \sqrt{l_j(n)}$ . Put

$$x_n = \prod_{1 \leq i \leq n-1} x_i^{(j)} + k, x_{n+1} = \prod_{1 \leq i \leq n-1} x_i^{(j)} + \frac{l_j(n)}{k}.$$

Direct computation shows that  $(x_1, x_2, \dots, x_{n+1})$  is a solution of Equation (4.33) $_{n+1}$ . The number of  $k$  such that  $1 < k < \sqrt{l_j(n)}$  is  $\frac{d(l_j(n))}{2} - 1$ . Therefore, we can construct  $\sum_{1 \leq j \leq \Omega(n-1)} (\frac{d(l_j(n))}{2} - 1)$  solutions of Equation (4.33) $_{n+1}$ , and all of the solutions satisfy  $x_{n+1} \neq \prod_{1 \leq i \leq n} x_i - 1$ .

**Step 2:** Let  $(x_1, x_2, \dots, x_n)$  be a solution of Equation (4.23) $_n$ . Put  $x_{n+1} = \prod_{1 \leq i \leq n} x_i - 1$ . Then  $(x_1, x_2, \dots, x_{n+1})$  is a solution of Equation (4.33) $_{n+1}$ . Since  $x_{n+1} = \prod_{1 \leq i \leq n} x_i - 1$ , this solution is different from any other solutions generated in step 1. Therefore,  $A(n+1) \geq \Omega(n) + \sum_{1 \leq j \leq \Omega(n-1)} (\frac{d(l_j(n))}{2} - 1)$ .  $\square$

**Lemma 4.5.16.** *For any solution  $(x_1, x_2, \dots, x_n)$  of Equation (4.23) $_n$ , put*

$$j = \prod_{1 \leq i \leq n} x_i, x_{n+1} = j + k, x_{n+2} = j + \frac{j^2 + t}{k}, x_{n+3} = \frac{1}{t} [j(j+k)(j + \frac{j^2 + t}{k}) - 1],$$

where  $k, t$  are positive integers. If  $x_{n+1} < x_{n+2} < x_{n+3}$  and  $x_{n+1}, x_{n+2}, x_{n+3} \in \mathbb{Z}$ , then  $(x_1, x_2, \dots, x_{n+3})$  is a solution of Equation (4.33) $_{n+3}$ , and

$$k|(j^2 + t), t|[j(j+k)(j + \frac{j^2+t}{k}) - 1], \gcd(j, k) = \gcd(k, t) = \gcd(j, t) = 1.$$

Moreover,  $2|j$  implies (i) the Jacobi symbol  $(\frac{k}{t}) = 1$ ; (ii)  $k \equiv 1 \pmod{4}$  or  $k \equiv t \equiv 3 \pmod{4}$ .

*Proof.* Assume that  $(x_1, x_2, \dots, x_n)$  is a solution of Equation (4.23) $_n$ .  $j = \prod_{1 \leq i \leq n} x_i$ ,  $x_{n+1} = j + k$ ,  $x_{n+2} = j + \frac{j^2+t}{k}$ ,  $x_{n+3} = \frac{1}{t}[j(j+k)(j + \frac{j^2+t}{k}) - 1]$ , where  $k$ , and  $t$  are positive integers. Assume that  $x_{n+1} < x_{n+2} < x_{n+3}$  and  $x_{n+1}, x_{n+2}, x_{n+3} \in \mathbb{Z}$ . Direct computation shows that  $(x_1, x_2, \dots, x_{n+3})$  is a solution of Equation (4.33) $_{n+3}$ . Put  $y = x_{n+1}$ ,  $z = x_{n+2}$ ,  $w = x_{n+3}$ . We have

$$\sum_{1 \leq i \leq n+3} \frac{1}{x_i} - \prod_{1 \leq i \leq n+3} \frac{1}{x_i} = 1,$$

$$1 - \frac{1}{j} + \frac{1}{y} + \frac{1}{z} + \frac{1}{w} - \frac{1}{jyzw} = 1.$$

Therefore,  $zwy + ywj + yzj - 1 = yzw$ .

Let  $d$  denote  $\gcd(j, k)$ , we have  $d|y$  since  $y = j + k$ . Therefore  $d|yzw$ , and so  $d|1$ .

Let  $d'$  denote  $\gcd(k, t)$ . Since  $t|[j(j+k)(j + \frac{j^2+t}{k}) - 1]$ , we have

$$d'|[j(j+k)(j + \frac{j^2+t}{k}) - 1].$$

From  $d'|k$ , we get

$$d'|[j^2(j + \frac{j^2+t}{k}) - 1], \quad k|(j^2 + t).$$

Therefore  $d'|j^2$ . We have  $d'|1$ .

Let  $d''$  denote  $\gcd(j, t)$ . From  $wt = j(j+k)(j + \frac{j^2+t}{k}) - 1$ , and  $d''|j$ ,  $d''|wt$ , it

follows that  $d'' = 1$ .

Hence,  $\gcd(j, k) = \gcd(k, t) = \gcd(j, t) = 1$ .

Assume that  $2|j$ . Since  $t|[j(j+k)(j+\frac{j^2+t}{k})-1]$ , we have  $2 \nmid t$ . From  $k|(j^2+t)$  we have  $2 \nmid k$ . Then from  $k|(j^2+t)$  we have

$$\left(\frac{-t}{k}\right) = 1, \quad j(j+k)(j+\frac{j^2+t}{k})-1 \equiv 0 \pmod{t}.$$

Hence  $j^2(j+k)^2 \equiv k \pmod{t}$ . We have

$$\left(\frac{k}{t}\right) = 1, \quad 1 = \left(\frac{-t}{k}\right)\left(\frac{k}{t}\right) = (-1)^{\frac{k-1}{2}}\left(\frac{t}{k}\right)\left(\frac{k}{t}\right) = (-1)^{\frac{k-1}{2} + \frac{(k-1)(t-1)}{4}}.$$

Hence,  $k \equiv 1 \pmod{4}$  or  $k \equiv t \equiv 3 \pmod{4}$ . □

**Lemma 4.5.17.**  *$\mathcal{A}$  is any positive integer. Put  $x_1 = \mathcal{A}$ . For  $1 < i \leq n$ , let  $x_i$  be the minimum positive integer satisfying  $x_i > x_{i-1}$  and  $x_i$  is coprime to all  $x_1, x_2, \dots, x_{i-1}$ . If  $\sum_{1 \leq i \leq n} \frac{1}{x_i} - \prod_{1 \leq i \leq n} \frac{1}{x_i} < 1$ , there is no solution of Equation (4.33)<sub>n</sub> with  $x_1 = \mathcal{A}$ .*

*Proof.* Suppose there is one solution of Equation (4.33)<sub>n</sub> with  $x_1 = \mathcal{A}$ , which is  $(\mathcal{A}, x_2^*, \dots, x_n^*)$ . For  $1 < i \leq n$ , since  $x_i$  is the minimum positive integer satisfying  $x_i > x_{i-1}$  and  $x_i$  is coprime to all  $x_1, x_2, \dots, x_{i-1}$ . We have

$$\begin{aligned} \sum_{1 \leq i \leq n} \frac{1}{x_i} - \prod_{1 \leq i \leq n} \frac{1}{x_i} &\geq \frac{1}{x_1^*} + \sum_{2 \leq i \leq n} \frac{1}{x_i} - \frac{1}{x_1^*} \prod_{2 \leq i \leq n} \frac{1}{x_i} \geq \\ \frac{1}{x_1^*} + \frac{1}{x_2^*} + \sum_{3 \leq i \leq n} \frac{1}{x_i} - \frac{1}{x_1^* x_2^*} \prod_{3 \leq i \leq n} \frac{1}{x_i} &\geq \dots \geq \sum_{1 \leq i \leq n} \frac{1}{x_i^*} - \prod_{1 \leq i \leq n} \frac{1}{x_i^*} = 1. \end{aligned}$$

This completes the proof. □

**Lemma 4.5.18.** *Suppose  $2 \leq x_1 < \dots < x_{n-2}$  and  $x_1, x_2, \dots, x_{n-2}$  are pairwise coprime. Put*

$$A = \prod_{1 \leq i \leq n-2} x_i, B = A - \sum_{1 \leq i \leq n-2} \frac{A}{x_i}.$$



Let  $P$  and  $Q$  be positive integers such that  $PQ = A^2 - B$ ,  $P < Q$ . Put

$$x_{n-1} = \frac{A+P}{B}, x_n = \frac{A+Q}{B}.$$

If  $x_{n-2} < x_{n-1}$ ,  $x_{n-1} \in \mathbb{Z}$  and  $x_n \in \mathbb{Z}$ , then  $(x_1, x_2, \dots, x_n)$  is a solution of Equation (4.33)<sub>n</sub>.

*Proof.* By a direct computation, it is easy to prove the lemma. □

### Algorithm Outline—Direct Finding Algorithms

Let  $(x_1, x_2, \dots, x_n)$  be a solution of Equation (4.33)<sub>n</sub>.

- First, we calculate  $L_1$  (the lower bound of  $x_1$ ) and  $U_1$  (the upper bound of  $x_1$ ).

Let  $x_1$  be  $\mathcal{A}$ , and let  $x_{i+1}$  be the minimum positive integer such that  $x_{i+1} > x_i$  and  $x_{i+1}$  are coprimes to any of  $x_1, x_2, \dots, x_i$ ,  $1 \leq i < n$ . If  $\sum_{1 \leq i \leq n} \frac{1}{x_i} - \prod_{1 \leq i \leq n} \frac{1}{x_i} \geq 1$ , then  $\mathcal{A} \in [L_1, U_1]$  by Lemma 4.5.17.

For example, for Equation (4.33)<sub>8</sub>, we fix  $\mathcal{A} = 2$ , then we get  $x_1 = 2, x_2 = 3, x_3 = 5, x_4 = 7, x_5 = 11, x_6 = 13, x_7 = 17$ , and  $x_8 = 19$ .

$$\frac{1}{2} + \frac{1}{3} + \frac{1}{5} + \frac{1}{7} + \frac{1}{11} + \frac{1}{13} + \frac{1}{17} + \frac{1}{19} - \frac{1}{2 \cdot 3 \cdot 5 \cdot 7 \cdot 11 \cdot 13 \cdot 17 \cdot 19} > 1.$$

We fix  $\mathcal{A} = 3$ , then we get  $x_1 = 3, x_2 = 4, x_3 = 5, x_4 = 7, x_5 = 11, x_6 = 13, x_7 = 17$ , and  $x_8 = 19$ .

$$\frac{1}{3} + \frac{1}{4} + \frac{1}{5} + \frac{1}{7} + \frac{1}{11} + \frac{1}{13} + \frac{1}{17} + \frac{1}{19} - \frac{1}{3 \cdot 4 \cdot 5 \cdot 7 \cdot 11 \cdot 13 \cdot 17 \cdot 19} > 1.$$

We fix  $\mathcal{A} = 4$ , then we get  $x_1 = 4, x_2 = 5, x_3 = 7, x_4 = 9, x_5 = 11, x_6 =$

$13, x_7 = 17$ , and  $x_8 = 19$ .

$$\frac{1}{4} + \frac{1}{5} + \frac{1}{7} + \frac{1}{9} + \frac{1}{11} + \frac{1}{13} + \frac{1}{17} + \frac{1}{19} - \frac{1}{4 \cdot 5 \cdot 7 \cdot 9 \cdot 11 \cdot 13 \cdot 17 \cdot 19} < 1.$$

Hence,  $L_1 = 2, U_1 = 3$ .

- Then for all values of  $x_1, x_2, \dots, x_i, x_j \in [L_j, U_j], 1 \leq j \leq i$ , we calculate  $L_{i+1}$  and  $U_{i+1}, 1 \leq i < n - 1$ .

– From

$$\frac{1}{x_1} + \dots + \frac{1}{x_i} < \sum_{1 \leq j \leq n} \frac{1}{x_j} - \prod_{1 \leq j \leq n} \frac{1}{x_j} = 1,$$

we have

$$x_i > \frac{\prod_{1 \leq k \leq i-1} x_k}{\prod_{1 \leq k \leq i-1} x_k - \sum_{1 \leq k \leq i-1} \prod_{1 \leq j \leq i-1, j \neq k} x_j}.$$

Hence,

$$L_i = \max \left\{ x_{i-1} + 1, \left\lceil \frac{\prod_{1 \leq k \leq i-1} x_k}{\prod_{1 \leq k \leq i-1} x_k - \sum_{1 \leq k \leq i-1} \prod_{1 \leq j \leq i-1, j \neq k} x_j} \right\rceil \right\}.$$

– It follows from

$$\sum_{1 \leq k < i} \frac{1}{x_k} + (n - i + 1) \frac{1}{x_i} > \sum_{1 \leq i \leq n} \frac{1}{x_i} - \prod_{1 \leq i \leq n} \frac{1}{x_i} = 1$$

that

$$x_i < \frac{\left( \prod_{1 \leq k \leq i-1} x_k \right) (n - i + 1)}{\prod_{1 \leq k \leq i-1} x_k - \sum_{1 \leq k \leq i-1} \prod_{1 \leq j \leq i-1, j \neq k} x_j}.$$

Therefore,

$$U_i = \left\lfloor \frac{\left( \prod_{1 \leq k \leq i-1} x_k \right) (n - i + 1)}{\prod_{1 \leq k \leq i-1} x_k - \sum_{1 \leq k \leq i-1} \prod_{1 \leq j \leq i-1, j \neq k} x_j} \right\rfloor.$$

- When  $(x_1, x_2, \dots, x_{i-1})$  are fixed, we set  $x_i = U_i$  and we let  $x_j$  ( $i < j \leq n$ ) be the minimum positive integer such that  $x_j$  is coprime to any of  $x_1, x_2, \dots, x_{j-1}$  and  $x_j > x_{j-1}$ . Then

$$\sum_{1 \leq i \leq n} \frac{1}{x_i} - \prod_{1 \leq i \leq n} \frac{1}{x_i} \geq 1.$$

We can reduce  $U_i$  based on this fact.

- When  $(x_1, x_2, \dots, x_{n-2})$  are fixed, calculate  $x_{n-1}, x_n$  by Lemma 4.5.18. Put

$$A = \prod_{1 \leq i \leq n-2} x_i, \quad B = A - \sum_{1 \leq i \leq n-2} \frac{A}{x_i}.$$

Factorize  $A^2 - B$ . For all positive integers  $P$  and  $Q$  such that  $PQ = A^2 - B$ ,  $P < Q$ , put

$$x_{n-1} = \frac{A + P}{B}, \quad x_n = \frac{A + Q}{B}.$$

If  $x_{n-2} < x_{n-1}$ ,  $x_{n-1} \in \mathbb{Z}$  and  $x_n \in \mathbb{Z}$ , then  $(x_1, x_2, \dots, x_n)$  is a solution of Equation (4.33)<sub>n</sub>.

## Pseudo Codes

The pseudo codes of the above algorithms are listed as follows:

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### Algorithm 3 Search2

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**Procedure** Search2(*depth*)

**Input:** *depth*

**Output:** All solutions of Equation (4.33)<sub>n</sub>

1: **if** *depth* = *n* **then**

2:      $x_n \leftarrow \frac{\prod_{1 \leq k \leq n-1} x_k - 1}{\prod_{1 \leq k \leq n-1} x_k - \sum_{1 \leq k \leq n-1} \prod_{1 \leq i \leq n-1, i \neq k} x_i}$

3:     **if**  $x_n$  is an integer **then**

```

4:         output  $x_1, x_2, \dots, x_n$ 
5:     end if
6: else
7:     if  $depth < n - 1$  then
8:          $L_{depth} \leftarrow \max \{x_{depth-1} + 1,$ 
9:              $\left\lceil \frac{\prod_{1 \leq k \leq depth-1} x_k}{\prod_{1 \leq k \leq depth-1} x_k - \sum_{1 \leq k \leq depth-1} \prod_{1 \leq j \leq depth-1, j \neq k} x_j} \right\rceil \}$ 
10:         $U_{depth} \leftarrow \left\lfloor \frac{(\prod_{1 \leq k \leq depth-1} x_k)(n - depth + 1)}{\prod_{1 \leq k \leq depth-1} x_k - \sum_{1 \leq k \leq depth-1} \prod_{1 \leq j \leq depth-1, j \neq k} x_j} \right\rfloor$ 
11:        call ReduceUpperBound( $depth$ ).
12:        for all  $x_{depth} \in [L_{depth}, U_{depth}]$  do
13:            if  $(x_1, \dots, x_{depth})$  are all relatively prime then
14:                call Search2( $depth + 1$ )
15:            end if
16:        end for
17:    else
18:        if  $depth = n - 1$  then
19:             $L_{depth} \leftarrow \max \{x_{depth-1} + 1,$ 
20:                 $\left\lceil \frac{\prod_{1 \leq k \leq depth-1} x_k}{\prod_{1 \leq k \leq depth-1} x_k - \sum_{1 \leq k \leq depth-1} \prod_{1 \leq j \leq depth-1, j \neq k} x_j} \right\rceil \}$ 
21:             $U_{depth} \leftarrow \left\lfloor \frac{(\prod_{1 \leq k \leq depth-1} x_k)(n - depth + 1)}{\prod_{1 \leq k \leq depth-1} x_k - \sum_{1 \leq k \leq depth-1} \prod_{1 \leq j \leq depth-1, j \neq k} x_j} \right\rfloor$ 
22:            call ReduceUpperBound( $depth$ )
23:            if  $U_{depth} - L_{depth} < 120000$  (see the remark), then
24:                for all  $x_{depth} \in [L_{depth}, U_{depth}]$  do
25:                    if  $(x_1, \dots, x_{depth})$  are all relatively prime then
26:                        call Search2( $depth + 1$ )
27:                    end if
28:                end for
29:            else
30:                 $A \leftarrow \prod_{1 \leq i \leq n-2} x_i$ 

```

```

29:       $B \leftarrow A - \sum_{1 \leq i \leq n-2} \frac{A}{x_i}$ 
30:       $C \leftarrow A^2 - B$ 
31:      for all positive factors  $P$  of  $C$  (see the remark) do
32:           $Q \leftarrow \frac{C}{P}$ 
33:          If  $P < Q$  then
34:               $x_{n-1} \leftarrow \frac{A+P}{B}, x_n \leftarrow \frac{A+Q}{B}$ 
35:              if  $x_{n-1} > x_{n-2}$  and both  $x_{n-1}$  and  $x_n$  are integers
                  then
36:                  return  $x_1, x_2, \dots, x_n$ 
37:              end if
38:          end if
39:      end for
40:  end if
41:  end if
42:  end if
43: end if

```

---



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#### Algorithm 4 Reduce Upper Bound

---

**Procedure:** ReduceUpperBound( $depth$ )

**Input:**  $depth$

**Output:** Nothing

```

1:  while TRUE do
2:       $x_{depth} \leftarrow U_{depth}$ 
3:      for  $i = depth + 1$  to  $n$  do
4:           $x_i \leftarrow$  the minimum value such that  $(x_1, x_2, \dots, x_i)$  are all relatively
                        prime
5:      end for

```

```

6:      if  $\sum_{1 \leq i \leq n} \frac{1}{x_i} - \prod_{1 \leq i \leq n} \frac{1}{x_i} < 1$  then
7:           $U_{depth} \leftarrow U_{depth} - 1$ 
8:      else
9:          return
10:     end if
11: end while

```

---

*Remark 4.5.19.* The complete solutions of Equation (4.33) are found by running Algorithm  $\text{Search2}(\text{depth} = 1)$ .

*Remark 4.5.20.* Algorithm *Search2* can be further optimized. The computation amount of factoring an integer is about 60,000 times as that of computing  $x_n$  from  $(x_1, x_2, \dots, x_{n-1})$ . Hence, we factorize  $A^2 - B$  to get  $x_{n-1}, x_n$  only if  $U_{n-1} - L_{n-1} > 120,000$ .

*Remark 4.5.21.* The integer factorization algorithm is based on Alpern's [4] contribution.

### New Solutions Generated from Known Solutions

The new solutions of Minus-type Equation (4.33) can be extracted from the solutions of Plus-type Equation (4.23). The following four methods generate new solutions from known solutions of Equation (4.23).

- **Method 1: Gap-one extension.** For each solution  $(x_1, x_2, \dots, x_n)$  of Equation (4.23)<sub>n</sub>, we set  $x_{n+1} = \prod_{1 \leq i \leq n} x_i - 1$ , then  $(x_1, x_2, \dots, x_{n+1})$  is a solution of Equation (4.33)<sub>n+1</sub>.
- **Method 2: Gap-two extension.** For each solution  $(x_1, x_2, \dots, x_n)$  of Equation (4.23)<sub>n</sub>, we set  $J = \prod_{1 \leq i \leq n} x_i$ ,  $C = J^2 - 1$ . Then we factor  $C$ . For

each positive factor  $k$  of  $C$ , we set

$$x_{n+1} = J + k, \quad x_{n+2} = J + \frac{C}{k}.$$

If  $x_n < x_{n+1} < x_{n+2}$ , then  $(x_1, x_2, \dots, x_{n+2})$  is a solution of Equation (4.33) $_{n+2}$ .

- Method 3: Gap-three extension. For each solution  $(x_1, x_2, \dots, x_n)$  of Equation (4.23) $_n$ , put  $J = \prod_{1 \leq i \leq n} x_i$ . For each odd integer  $t > 0$ , one factors  $J^2 + t$ . For each positive factor  $k$  of  $J^2 + t$ , put

$$x_{n+1} = J + k, \quad x_{n+2} = J + \frac{J^2 + t}{k}, \quad x_{n+3} = \frac{1}{t}[J \cdot x_{n+1} \cdot x_{n+2} - 1].$$

Then  $(x_1, x_2, \dots, x_{n+3})$  is probably a solution of Equation (4.33) $_{n+3}$ .

- Method 4: Gap-two extension<sup>+</sup>. Randomly choose  $(x_1, x_2, \dots, x_n)$ , such that  $x_1, x_2, \dots, x_n$  are pairwise coprime,  $2 \leq x_1 < \dots < x_n$ . Let

$$A = \prod_{1 \leq i \leq n} x_i, \quad B = A - \sum_{1 \leq i \leq n} \frac{A}{x_i}.$$

For all  $P, Q$  such that  $1 \leq P < Q$ ,  $PQ = A^2 - B$ , let

$$x_{n+1} = \frac{A + P}{B}, \quad x_{n+2} = \frac{A + Q}{B}.$$

If  $x_n < x_{n+1}$ ,  $x_{n+1} \in \mathbb{Z}$  and  $x_{n+2} \in \mathbb{Z}$ , then  $(x_1, x_2, \dots, x_{n+2})$  is a solution of Equation (4.33) $_{n+2}$ .

**The number of solutions for  $n = 8, 9, 10$**

All the solutions of Minus-Type Equations for  $n \leq 7$  are listed in Appendix B. We also have the following results:

**Theorem 4.5.22.** *The number of the complete solutions of Equation (4.33)<sub>8</sub> is 550 (listed in Appendix C).*

**Corollary 4.5.23.** *There are only 12 Giuga numbers with no more than eight distinct prime factors.*

*Proof.* A number  $K$  is a Giuga number if and only if  $\sum_{p|K} \frac{1}{p} - \prod_{p|K} \frac{1}{p} \in \mathbb{N}$  [124]. If  $n$  has no more than 8 distinct prime factors, then  $K$  is a Giuga number if and only if  $\sum_{p|K} \frac{1}{p} - \prod_{p|K} \frac{1}{p} = 1$ , since

$$\sum_{p|K} \frac{1}{p} - \prod_{p|K} \frac{1}{p} \leq \frac{1}{2} + \frac{1}{3} + \frac{1}{5} + \frac{1}{7} + \frac{1}{11} + \frac{1}{13} + \frac{1}{17} + \frac{1}{19} < 2.$$

Each solution of Equation (4.33) <sub>$n$</sub>  in which  $x_1, x_2, \dots, x_n$  are all primes corresponding to a Giuga number  $\prod_{1 \leq i \leq n} x_i$ . Therefore, this corollary can be verified by all solutions of Equation (4.33) <sub>$n$</sub>  for  $n \leq 8$ , which are listed in Appendices B and C. Giuga numbers with no more than eight distinct prime factors are:

$$2 \times 3 \times 5,$$

$$2 \times 3 \times 7 \times 41,$$

$$2 \times 3 \times 11 \times 13,$$

$$2 \times 3 \times 11 \times 17 \times 59,$$

$$2 \times 3 \times 7 \times 43 \times 3041 \times 4447,$$

$$2 \times 3 \times 11 \times 23 \times 31 \times 47057,$$

$$2 \times 3 \times 7 \times 59 \times 163 \times 1381 \times 775807,$$

$$2 \times 3 \times 7 \times 71 \times 103 \times 61559 \times 29133437,$$

$$2 \times 3 \times 7 \times 71 \times 103 \times 67213 \times 713863,$$

$$2 \times 3 \times 7 \times 43 \times 1831 \times 138683 \times 2861051 \times 1456230512169437,$$

$$2 \times 3 \times 11 \times 23 \times 31 \times 47059 \times 2259696349 \times 110725121051,$$

$$2 \times 3 \times 11 \times 23 \times 31 \times 47137 \times 28282147 \times 3892535183.$$

□



**Theorem 4.5.24.** *Equation  $(4.33)_9$  has at least 1547 solutions. These solutions are available at <http://tdt.sjtu.edu.cn/9minus.html>*

*Proof.* The number of solutions found by each method are listed in Table 4.12.

Table 4.12: Numbers of solutions of Equation  $(4.33)_9$  found

Numbers of solutions found	Method
122	Method 1
1342	Method 2
51	Method 4
32	Direct finding

□

**Theorem 4.5.25.** *Equation  $(4.33)_{10}$  has at least 18984 solutions. These solutions are available at <http://tdt.sjtu.edu.cn/10minus.html>*

*Proof.* From the known 411 solutions of Equation  $(4.23)_9$ , we can find 411 solutions of Equation  $(4.33)_{10}$  by the Method 1 (Gap-one extension). The others are from the Method 2 (Gap-two extension). □

## 4.6 Notes

In this chapter, we have introduced a new direction of modern cryptography—*Batch Cryptography*, which includes the batch encryption/decryption, batch key agreement, and batch signature/verification. We implemented batch RSA schemes based on the Diophantine Equations (Plus-Type Equations and Minus-Type Equations), and gave the methods to solve these Diophantine equations<sup>2</sup>.

There are some interesting problems in batch cryptography which are still open. For instance,

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<sup>2</sup>We have finished two manuscripts about efficiently implementing batch RSA and solving the Diophantine Equations (Plus-Type Equations and Minus-Type Equations) in 2008. The readers who are interested in this area can refer to [71, 72].

1. How to design aggregate signature and batch verification schemes without extra requirements, such that they are secure in the standard model?

*Researches in the batch cryptography are focusing on the aggregate signature and batch verification, which are secure in the random oracle model. In order to make them secure in the standard model, it is usual to introduce extra requirements. However, it remains open to construct a practical aggregation scheme in the standard model without extra requirements such as timing, interactive restrictions, or requiring each user to be able to prove knowledge of his/her secret key. It is also open to explore other relaxations of the full aggregation model.*

2. How to design batch encryption schemes based on other assumptions such as quadratic residue, discrete logarithm, and pairing based assumptions?

*Almost all existing researches are focusing on the batch RSA. To avoid putting all eggs in one basket, it is better to design new feasible batch cryptographic algorithms based on quadratic residues or discrete logarithm problem. Besides, there is no formal security proof for the batch RSA. This is still an open problem.*

3. How to design batch key exchange/agreement protocols?

*So far, the batch key exchange/agreement has attracted little attention. In the special case where there are many senders and only one receiver, each sender sends the ciphertext of a random number to the receiver under the receiver's public key. Then, the receiver is able to batch decrypt them and establish a secure channel with each sender. To design batch key exchange/agreement protocols becomes more complicated if the attackers exist and the property of entity authentication is required.*

4. How to combine batch technique with other cryptographic algorithms?

*In this chapter, we only considered batch algorithms without combining them with other algorithms, such as proxy re-encryption/re-signature and attribute-based encryption/signature. It must be interesting to construct cryptographic algorithms using batch techniques to improve the efficiency as well as the security.*

5. How to find all the solutions of the Plus-Type Equations and Minus-Type Equations with  $n = 9, 10$ , which can be used as the public keys in batch RSA? Furthermore, how to find the solutions to these equations when the right side of these equations is 2, not 1?

*We have given the new implementation of batch RSA. It requires only one modular,  $N$ . This implementation is more suitable to construct hardware circuits for batch RSA decryption. This implementation is based on the Diophantine Equations  $\sum_{1 \leq i \leq N} \frac{1}{x_i} \pm \prod_{1 \leq i \leq N} \frac{1}{x_i} = 1$ ,  $2 \leq x_1 < x_2 < \dots < x_N$ . All solutions to these equations for  $N \leq 8$  are known, but the complete solutions for  $N \geq 9$  are unknown.*

## Chapter 5

# Noncommutative Cryptography

### 5.1 Introduction

As a new direction in modern cryptography, noncommutative cryptography differs much from those discussed in previous chapters. In this chapter, we will discuss this subject through exploring the following questions.

*Question 1: What is noncommutative cryptography?*

An immediate reply: It is an abbreviation of the cryptography based on noncommutative algebraic structures, such as

- noncommutative groups or semigroups,
- noncommutative rings,
- even arbitrary sets over which some noncommutative operations could be well-defined.

*Question 2: Why do we need noncommutative cryptography?*

Before answering this question, let us pay attention to the following issues:

- In 1994, Shor [257] proposed an efficient quantum algorithm for solving integer factoring problem (IFP<sup>1</sup>) and the discrete logarithm problem (DLP). In 2003, Shor's algorithm was also extended to solve the discrete logarithm problem over elliptic curves (ECDLP) [229]. Shor's algorithm can be regarded as a special case of Kitaev's framework [172] for solving the so-called hidden subgroup (or subfield) problems (HSP). Now, we are aware of efficient quantum algorithms for HSP over arbitrary *commutative* groups. But there are evidences to suggest that HSP over *noncommutative* groups might be much harder: the progress in quantum algorithms for HSP over some noncommutative groups (e.g., symmetric groups) is very limited, even negative in some cases [234].
- In 2002, Stinson et al. [207] observed that most unbroken public-key cryptosystems used today were based on commutative algebraic structures, such as
  - $\mathbb{Z}_N$  (where  $N$  is a large composite integer),  $\mathbb{Z}_p$  (where  $p$  is a large prime), and  $\mathbb{F}_q$  (where  $q = p^m$  is a power of some prime  $p$  and  $m$  is a positive integer), over which the well-known cryptosystems such as RSA encryption, ElGamal encryption, and Schnorr signature, are defined respectively;
  - $\mathbb{E}(\mathbb{F}_q)$  (i.e., the elliptic curves over the finite field  $F_q$ ) over which the elliptic curve cryptosystems and the pairing-based cryptosystems are defined;
  - $\mathbb{C}_K$  (i.e., the ideal class group over some algebraic (closed) field  $K$ ) over which the (real or imaginary) quadratic field cryptosystems are defined.

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<sup>1</sup>Shor's factoring algorithm uses two quantum registers. By introducing a more quantum register, we show that the measured numbers in the second and third quantum registers, which are of the same pre-measurement state, should be equal if Shor's complexity analysis is sound. This seems to contradict the Shor's argument that there are  $r$  possible observed values for the second register. Someone argues that the three quantum registers are entangled. If so, it is a peculiar entanglement which has not yet been mentioned. In our opinion, Shor's algorithm is not doubtless from a theoretical point of view. See Zhengjun Cao and Zhenfu Cao: On Shor's factoring algorithm with three quantum registers (under review).

For clarity, we would like to refer to these cryptosystems as *commutative cryptography*. The theoretical foundations for the above cryptosystems are based on the intractability of problems closer to number theory than group theory [207]. (A less rigid illustration is given in Figure 5.1.)

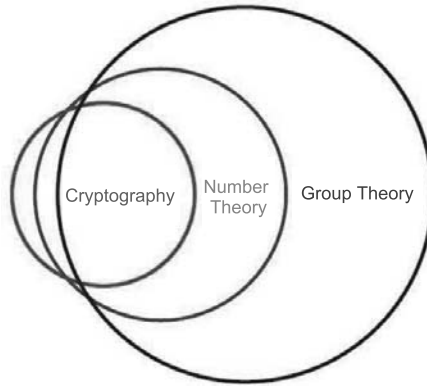


Figure 5.1: Cryptography, Number Theory, and Group Theory

From the above facts, we see that the classical commutative cryptography might be vulnerable to quantum computing. Moreover, cryptographers Goldreich and Lee [179] advised that *not to put all eggs in one basket*. Therefore, it seems applicable to introduce noncommutativity into cryptography.

*Question 3: How about the state of arts of noncommutative cryptography?*

The beginning of using noncommutative groups in cryptography goes back to Magyarik and Wagner [279] who proposed an approach to design public-key cryptosystems based on the undecidable word problem over groups and semigroups. This work has not attracted much attention until recently: In 2006, Birget et al. [28] pointed out that Magyarik and Wagner's scheme is actually not based on word problem, instead on premise problem, which is generally easier than the former. Now, when we look backward, it

seems that Magarik and Wagner's pioneering idea comes *a bit earlier*: The glory of the public-key cryptography community in 1980s was dominated by RSA, ElGamal, ECC etc. However, there is no silver bullet in this world. In 1994, Shor's efficient quantum algorithm for solving IFP and DLP emerged. Although practical quantum computer might be at least decades away, its potential powerful capability has already casted distrust in current cryptographic methods [179]. Since then many attempts have been made for developing alternative public-key cryptosystems based on different platforms. Therefore, noncommutative cryptography takes advantage of a turn of events and gets on.

In 1999, Anshel et al. [6] proposed an elegant algebraic key establishment protocol. The foundation of their method lies in the difficulty of solving equations over algebraic structures, in particular Garside groups [6]. In their pioneering paper, they also suggested that braid groups might be a good alternative. Shortly afterward, Ko et al. [175] published a fully fledged encryption scheme using braid groups. Since then the subject has met with a quick success (see Section 5.2). However, from 2001 to 2003, repeated cryptanalytic successes also diminished the initial optimism on the subject [93]. It seems that more researches are still needed to reach a definite conclusion on cryptographic potential of braid groups. Garber [117] made a comprehensive survey of braid-based cryptography in 2007.

In 2001, Paeng et al. [224] proposed another new public-key cryptosystem built on finite non-abelian groups. Paeng's method is based on the discrete logarithm problem in the inner automorphism group defined by the conjugate action. Paeng's systems were later improved and referred to as the so-called MOR system. In 2002, Magliveras et al. [207] also developed new approaches to design public key cryptosystems using one-way functions and trapdoors in finite groups. Their method originated from group theory. They introduced two public-key cryptosystems, MST1 and MST2, based on the difficulty of computing certain factorizations in finite groups. Later, Vasco et al. [276]

demonstrated that the factorization concept used in MST1 and MST2 admits a uniform description of several cryptographic primitives. Also, it turned out that a generalization of MST2 can serve as a unified framework for several public-key cryptosystems, including the ElGamal system [107], the braid group based system [175] and the MOR system.

In 2002, Grigoriev and Ponomarenko firstly constructed [131] homomorphic cryptosystems over non-abelian groups. Shortly afterward, based on the difficulty of the membership problem for groups of integer matrices, their method was extended to arbitrary finite groups [132, 133].

In 2004, inspired by Anshel et al.'s idea in the algebraic key exchange, Eick and Kahrobaei [106] proposed a new cryptosystem over *polycyclic groups*<sup>2</sup>. Further, Mahalanobis [208] in 2006 generalized the Diffie–Hellman key exchange protocol from a cyclic group to a finitely presented non-abelian *nilpotent group* of class 2.

In 2005, Shpilrain and Ushakov [259] suggested that Thompson's group might be a good platform for constructing public-key cryptosystems. In their contribution, the basic assumption is the intractability of the decomposition problem, which is more general than the conjugator search problem.

In 2006, Baumslag et al. [19] suggested potential cryptosystems using linear groups. In 2007, they [17] further suggested to use the classical modular group as a platform for cryptography. In fact, Yamamura [299] should be credit for introducing the extended modular group  $SL_2(\mathbb{Z})$  into cryptography in 1998. But Yamamura's scheme was shown to have loopholes [135]. In [17] attacks based on these loopholes were closed [112].

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<sup>2</sup>In group theory, a *normal series* is a series of normal subgroups of a group  $G$ ,

$$G = H_1 \triangleright H_2 \triangleright \dots \triangleright H_{n+1} = \{1\}.$$

If each term of the series is normal not only in the whole group but also in the preceding term, then the series is called *subnormal*. A group  $G$  is called *polycyclic* if it has a subnormal series with cyclic factors, i.e.,  $H_i/H_{i+1}$  is cyclic for  $i = 1, \dots, n$ . A group  $G$  is called *nilpotent* if it has a normal series so that every quotient  $H_i/H_{i+1}$  lies in the centre of  $G/H_{i+1}$  (a so-called central series). The length of a shortest central series of a nilpotent group is called its class (or degree of nilpotency). Finitely-generated nilpotent groups are polycyclic groups and, moreover, have a central series with cyclic factors.



Also in 2006, Dehornoy [94] proposed an authentication scheme based on the left self-distributive (LD) systems. This idea was further developed by introducing the concept of the one-way LD system [287]. An algebraic system  $(A, *)$  is called *left self-distributive* if for all  $a, b, c \in A$ ,

$$a * (b * c) = (a * b) * (a * c). \quad (5.1)$$

An LD system is said one-way if it is intractable to extract  $a$  for giving  $a * b$  and  $b$ . In general, an LD system is much different from (semi) groups. In fact, it is even not associative. However, one can easily define non-trivial LD systems over any noncommutative group  $G$  via the mapping

$$a * b \triangleq aba^{-1}. \quad (5.2)$$

Moreover, if the conjugator search problem (CSP) over  $G$  is intractable, then the derived LD system is one-way. Therefore, cryptosystems over one-way LD systems admit many kinds of implementations.

In 2006, we [73] proposed a method to use polynomials over noncommutative rings or (semi) groups to build cryptographic schemes. This method is now referred to as the  $\mathbb{Z}$ -modular method (See Section 5.3). In 2008,  $\mathbb{Z}$ -modular method was used to build signature schemes over noncommutative groups and division *semirings*<sup>3</sup> [231], respectively. Moreover, if we restrict the polynomials used in  $\mathbb{Z}$ -modular method to be monomials, then the conjugacy related assumptions can be viewed as special cases of the assumptions defined in  $\mathbb{Z}$ -modular method.

In 2009, Vats [277] proposed a cryptosystem, named as NNRU, and claimed that it is a noncommutative analogue of the well-known NTRU cryptosystem [143]. NNRU

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<sup>3</sup>A semiring is a natural noncommutative generalization of a ring in the sense that in both the binary operations  $+$  and  $\cdot$  are not required to be commutative.

operates in the noncommutative ring  $\mathbf{M} = M_k(\mathbb{Z}[X]/(X^n - I_{k \times k}))$ , where  $\mathbf{M}$  is a matrix ring of  $k \times k$  matrices of polynomials in  $R = \mathbb{Z}[X]/(X^n - 1)$ . The main enhancement is that the lattice-based attack which is the biggest threat to NTRU fail for NNRU.

In 2010, finite noncommutative groups of the four-dimension vectors over the ground field were constructed for implementing the cryptographic protocols [213].

In 2011, inspired by the recent success of applying the problems of learning parity with noise (LPN) [30] and learning with errors (LWE) [232] to a variety of cryptographic constructions, Baumslag et al. [18] proposed the generalized learning problem over noncommutative groups. Their work opens a new avenue for developing cryptography based on combinatorial group theory.

Of course, the above list is far from complete. Here, we just pick one or two typical examples in each year to illustrate that the research on noncommutative cryptography during the first decade of the new century is considerably attractive. We would like to point out that almost all initial attempts mentioned above have been shown to be insecure, but the experiences from these early attempts have a great significance. Readers who want to pay more attention to various attacks against these attempts and their security can refer to the recently published book *Group-based Cryptography (by Mayasnikov, Shpilrain, and Ushakov)* [218].

The rest of this chapter is organized as follows: In Section 5.2, we give a brief introduction on the well-known braid-based cryptography, including related fundamentals, cryptographic assumptions and related constructions, as well as our own contributions; In Section 5.3, we describe the  $\mathbb{Z}$ -modular method that is used to develop cryptographic schemes based on noncommutative rings or (semi) groups; In Section 5.4, by using monomials, instead of polynomials, we develop several cryptosystems based on intractability of the conjugator search problem and related assumptions; In Section 5.5, an improved key agreement protocol over Thompson's group is presented; And finally,

further remarks are addressed in Section 5.6.

## 5.2 Braid-Based Cryptography

### 5.2.1 Basic Definitions

For index  $n \geq 2$ , the braid group  $B_n$  is defined by the generators  $\sigma_1, \sigma_2, \dots, \sigma_{n-1}$  and the relations  $\sigma_i \sigma_j = \sigma_j \sigma_i$  for  $|i - j| > 1$  and  $\sigma_i \sigma_j \sigma_i = \sigma_j \sigma_i \sigma_j$  for  $|i - j| = 1$  ( $1 \leq i, j \leq n - 1$ ). This definition is called the *Artin presentation* and the generators are called *Artin's generators*. Intuitively, a geometrical illustration of the identity, denoted by  $e$ , of the braid group  $B_4$  and the Artin generators (e.g.,  $\sigma_2^{\pm 1}$ ) can be shown as Figure 5.2. Geometrically, the product of two braids is the braid obtained by merging the tail of the first braid with the head of the second braid. For example, Figure 5.3 shows the braid  $\sigma_1 \sigma_2 \sigma_1^{-1} \sigma_3^{-1} \sigma_2 \sigma_3 \sigma_2 \sigma_3^{-1} \sigma_2^{-1} \sigma_1$ . From the definition, we know that there is a natural automorphism from  $B_2$  to the integer additive group  $\mathbb{Z}$ . For  $n \geq 3$ , the braid group  $B_n$  is infinite and *noncommutative*. For each  $m (\leq n)$ , the identity mapping on  $\{\sigma_1, \dots, \sigma_{m-1}\}$  naturally induces an embedding of  $B_m$  into  $B_n$ . By this approach, it is easy to define the limit group  $B_\infty$ .

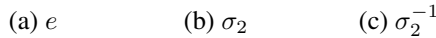


Figure 5.2: Geometrical illustration on identity and Artin generators

The relations used in the above definition shine the lights of topology on braids, that is, continuously moving the curves with their ends fixed does not change the value of a braid. But algebraically, having a *normal form* (i.e., a *unique presentation*) for each braid is very useful, since it lets us compare two braids without resorting on topological



Figure 5.3: An example of geometric braids

images. For braid groups, there are some known normal forms, one of which, *the (left) canonical form*, is commonly used in cryptography. With respect to the Artin presentation, the complexity of transforming a braid into its canonical form is  $\mathcal{O}(|w|^2 n \log n)$  where  $|w|$  is the length of  $w$ , i.e., the number of the Artin generators in  $w$  [26].

### 5.2.2 Conjugacy and Related Problems

For arbitrary two braids  $x, y \in B_n$ , we say that they are *conjugate*, written as  $x \sim y$ , if  $y = a^{-1}xa$  for some  $a \in B_n$ . Here  $a$  or  $a^{-1}$  is called a *conjugator*. In the braid group  $B_n$ , we can define the following cryptographic problems that are related to conjugacy [173, 286]:

- Conjugacy deciding problem (CDP): Determine whether  $x \sim y$  for a given instance  $(x, y) \in B_n^2$ .
- Conjugator searching problem (CSP): Find a braid  $z \in B_n$  so that  $y = z^{-1}xz$  for a given instance  $(x, y) \in B_n^2$  with  $x \sim y$ .
- Matched conjugate searching problem (MCSP): Find a braid  $y' \in B_n$  so that  $y \sim y'$  and  $xy \sim x'y'$  for a given instance  $(x, x', y) \in B_n^3$  with  $x \sim x'$ .
- Matched triple searching problem (MTSP): Find a triple  $(\alpha, \beta, \gamma) \in B_n^3$  so that  $\alpha \sim x, \beta \sim y, \alpha\beta \sim xy, \gamma \sim y$  and  $\alpha\gamma \sim x'y$  for a given instance  $(x, x', y) \in B_n^3$  with  $x \sim x'$ .

- *k*-simultaneous conjugator searching problem (*k*-SCSP): Given *k* pairs  $(x_1, x'_1), \dots, (x_k, x'_k) \in B_n^2$  with  $x'_i = s^{-1}x_i s$  for some  $s \in B_n$  and  $1 \leq i \leq k$ , find a braid  $z \in B_n$  so that  $x'_i = z^{-1}x_i z$  for  $1 \leq i \leq k$ .
- Decomposition problem (DP): Given  $(x, y) \in B_n^2$  and  $S_1, S_2 \subseteq B_n$ , find  $z_1 \in S_1$  and  $z_2 \in S_2$  so that  $y = z_1 x z_2$ .
- Root problem (RP): Given a positive integer *m* and a braid  $p \in B_n$  so that  $p$  is an *m*th power in  $B_n$ , find an *m*th root of  $p$ , i.e., find a braid  $r$  satisfying  $r^m = p$ .

By employing the terminologies used in [173], we say that a problem is

- *solvable*, if there is a deterministic finite algorithm that outputs an accurate solution.
- *unsolvable*, if it is not solvable.
- *tractable*, if there is a probabilistic polynomial-time algorithm that outputs an accurate solution with non-negligible probability (with respect to the length of description of the input instances).
- *intractable*, if it is not tractable.

At present, we know that all the above problems are solvable. However, we do not know whether they are tractable in general cases. It seems that all of them are intractable in the worst cases. According to [173, 286], we always take the following reduction relations into consideration:

$$CDP \prec MCSP \simeq MTSP \preceq CSP, \quad (5.3)$$

where  $\prec$ ,  $\preceq$ , and  $\simeq$  denote the relations of “easier than,” “not harder than,” and “as hard as,” respectively.

In [173], Ko et al. also described an efficient algorithm to solve CDP with *overwhelming accuracy*<sup>4</sup>. This algorithm is the common basis for all existing braid-based signatures because the verification algorithms need to determine whether two given braids are conjugate.

Also, MCSP *seems* as hard as CSP according to the heuristic analysis in [173]. In our opinion, this is why MCSP is used to design signature schemes in [173].

At present, the relationship between  $k$ -SCSP and CSP is unclear. Although we have not proven rigorously which of them is easier, some approximate methods and heuristic analysis suggest that the former *seems* easier than the latter [286]. Therefore, if a cryptographic scheme lays its security on the assumption of the intractability of  $k$ -SCSP, we say that it is suffered from the weakness of  $k$ -simultaneous conjugacy.

The relation between CSP and DP is also unclear at present. On one hand, a solution of CSP is also a solution for DP *only if*  $z^{-1} \in S_1$  and  $z \in S$  hold simultaneously; On the other hand, a solution for DP becomes a solution for CSP *only if*  $z_1 = z_2^{-1}$  holds.

According to [260], the only known algorithm for RP consists of explicitly enumerating several conjugacy classes related with the initial braid  $p$ . This enumerating process is exponential in essence and therefore becomes infeasible when the lengths of the braids are large enough. In practice, RP appears as even more difficult than CSP. Note that the braid groups are torsion-free, i.e., if  $b$  is a nontrivial braid, then  $b^m$  for every  $m \geq 2$ , is not trivial.

Although some algorithms for solving CSP were proposed [108, 113, 118, 125], none of them has been proven implementable with a polynomial-time complexity (with respect to the braid index  $n$ ). Until now, Gebhardt's algorithm [118], which was proposed in 2003 but formally published in 2005, is the most efficient method for solving CSP in braid groups. This algorithm has not yet been proven implementable within polynomial time. Subsequently, CSP in braid groups is classified for further study. According

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<sup>4</sup>We say an algorithm has *overwhelming accuracy* if its outputs are accurate with the probabilities that are negligibly close to 1.

to Garber's report [117], within polynomial time, we can only solve CSP in periodic braids. There exist two kinds of challenges in obtaining ultimate solutions for CSP in braid groups: one is how to solve CSP in rigid braids within polynomial time, and the other is how to find polynomial boundaries of the complexity of Gebhardt's method.

Considering that the assumption of intractability of CSP plays a central role in braid-based cryptography, Ko et al. [173] introduced the concept of *CSP-hard pair*. Let  $S_1$  and  $S_2$  be two subgroups of  $B_n$ . A pair  $(x, x') \in S_1 \times S_2$  is said *CSP-hard* if  $x \sim x'$  and this CSP instance is intractable. It is more meaningful to define a special sampling algorithm, *CSP-hard pair generator*, rather than to define a particular CSP-hard pair. CSP-hard pair generator  $\mathcal{K}_{csp}$  can be defined as follows:

- $\mathcal{K}_{csp}(n)$  is a probabilistic polynomial-time algorithm that takes the security parameter  $n$  as input, and outputs a triple  $(p, q, w) \in B_n^3$  so that  $q = w^{-1}pw$  holds and the CSP instance  $(p, q)$  is intractable.

For details regarding the methods to construct such a CSP-hard pair generator, please refer to [93], [173], [206] and [174]. In particular in [174], Ko et al. proposed several promising ways to generate hard instances of the conjugacy problems for braid cryptography.

### 5.2.3 Key Exchange, Encryption and Authentication

Now, let us have a quick review of the cryptographic protocols from braid groups. The first protocol was proposed theoretically by Anshel et al. in arbitrary Garside groups in 1999 [6], and implemented in braid groups in 2001 [5]. This protocol assumed that the CSP problem is difficult enough.

Suppose  $\Lambda = \{l_1, \dots, l_k\}$  be an alphabet. For a given word  $u$  on  $\Lambda$ , let  $u(p_1, \dots, p_k)$  denote the substitution of each  $l_i$  in  $u$  by  $p_i$  (for all  $1 \leq i \leq k$ ). Suppose that Alice's secret key is a word  $u$  on an alphabet of size  $k$ , while Bob's secret key is a

word  $v$  on a different alphabet of size  $m$ . The key exchange protocol works as follows:

**Protocol 5.2.1.** Anshel et al.'s Key Exchange:

- Public settings:  $p_1, \dots, p_k$  and  $q_1, \dots, q_m$  in  $B_n$ .

Private keys: Alice:  $u$ ; Bob:  $v$ .

- Exchange: Alice computes  $s = u(p_1, \dots, p_k)$ , and sends Bob the conjugates  $q'_1 = sq_1s^{-1}, \dots, q'_m = sq_ms^{-1}$ ; meanwhile, Bob computes  $r = v(q_1, \dots, q_m)$ , and sends Alice the conjugates  $p'_1 = rp_1r^{-1}, \dots, p'_k = rp_kr^{-1}$ .
- Key deriving: Alice computes  $K_A = H(su(p'_1, \dots, p'_k)^{-1})$ , and Bob computes  $K_B = H(v(q'_1, \dots, q'_m)r^{-1})$ , where  $H$  is the colored Burau representation defined by Morton [215]. (Essentially,  $H$  is a key deriving function and can be instantiated by proper hashes.)

After the execution of this protocol, both Alice and Bob share the same key  $K_A = K_B$ , since the following equalities hold:

$$\begin{aligned}
 su(p'_1, \dots, p'_k)^{-1} &= sru(p_1, \dots, p_k)^{-1}r^{-1} \\
 &= sr s^{-1}r^{-1} \\
 &= sv(q_1, \dots, q_m)s^{-1}r^{-1} \\
 &= v(q'_1, \dots, q'_m)r^{-1}.
 \end{aligned}$$

The security of this protocol is based on the difficulty of a variant of CSP in  $B_n$ , namely  $k$ -SCSP (see the above subsection). In [5], it is suggested to work in  $B_{80}$  with  $k = m = 20$  and short initial braids  $p_i, q_j$  with length 5 or 10.

The second protocol is proposed by Ko et al. [175]. Let  $LB_n$  (resp.  $UB_n$ ), a subgroup of  $B_n$ , be generated by  $\sigma_1, \dots, \sigma_{m-1}$  (resp.  $\sigma_{m+1}, \dots, \sigma_{n-1}$ ) with  $m = \lfloor \frac{n}{2} \rfloor$ . Then, every braid in  $LB_n$  commutes with every braid in  $UB_n$ .



**Protocol 5.2.2.** Ko et al.'s Key Exchange:

- Public settings: a braid  $p$  in  $B_n$ .

Private keys: Alice:  $s \in LB_n$ ; Bob:  $r \in UB_n$ .

- Exchange: Alice sends Bob  $p' = sps^{-1}$ ; meanwhile, Bob sends Alice  $p'' = rpr^{-1}$ .
- Key deriving: Alice computes  $K_A = sp''s^{-1}$ ; Bob computes  $K_B = rp'r^{-1}$ .

After the execution of this protocol, both Alice and Bob share the same key  $K_A = K_B = srpr^{-1}s^{-1}$ , since  $s$  and  $r$  commute. The security of this protocol is based on the difficulty of the so-called Diffie–Hellman like conjugacy problem (DHCP). DHCP is to find the braid  $rp'r^{-1}$ , or equivalently,  $sp''s^{-1}$  for given braids  $p, p', p'' \in B_n$ , where  $p' = sps^{-1}$  and  $p'' = rpr^{-1}$  for some  $s \in LB_n$  and  $r \in UB_n$ . The suggested parameters are  $n = 80$ , and the working braids are specified by normal sequences of length 12.

Furthermore, in the same paper [175], Ko et al. gave a fully fledged public-key encryption scheme based on the above Diffie–Hellman like key exchange protocol. Suppose a hash function  $h$ , which maps a braid into a bit-string with proper length, is at hand, and Alice's public key is the pair  $(p, p') \in B_n^2$  with  $p' = sps^{-1}$ , where  $s \in LB_n$  is Alice's private key. For sending a message  $m$  to Alice, Bob chooses a random braid  $r \in UB_n$  and sends the ciphertext  $c = m \oplus h(rp'r^{-1})$ , together with the additional datum  $p'' = rpr^{-1}$ . Now, Alice computes the plaintext  $m = c \oplus h(sp''s^{-1})$ . Apparently, this scheme is consistent and its security is also based on the difficulty of the DHCP problem. (Unfortunately, a deterministic polynomial time algorithm for solving the DHCP problem over braid groups was found by Cha et al. [82] in 2003. Thus, Ko et al.'s key exchange protocol, as well as the derived encryption scheme, is no longer secure.)

In 2006, Sibert et al. [260] designed three authentication schemes by using braid

word reduction. The first scheme is based on DHCP, the second is based on CSP, and the third is based on CSP and RP. Considering that the first becomes insecure<sup>5</sup> according to the aforementioned Cha et al.'s polynomial-time algorithm, we merely introduce the second and the third herein.

**Protocol 5.2.3.** Sibert et al.'s Authentication Scheme II:

- Key generation:
  - A(lice) chooses a public braid  $b \in B_n$ , so that the CSP problem for  $b$  is hard;
  - A chooses a secret braid  $s \in B_n$  as her private key; she computes  $b' = sbs^{-1}$  and publishes the pair  $(b, b')$  as her public key.
- Authentication: Repeat the following exchanges  $k$  times, where  $k$  is a polynomial function of the size of the braid specified:
  - A chooses a random braid  $r$ , and sends  $x = rbr^{-1}$  to B(ob);
  - B sends a random bit  $\epsilon$  to A;
  - For  $\epsilon = 0$ , A sends  $y = r$  to B, and B checks  $x = yby^{-1}$ ;
  - For  $\epsilon = 1$ , A sends  $y = rs^{-1}$  to B, and B checks  $x = yb'y^{-1}$ .

**Protocol 5.2.4.** Sibert et al.'s Authentication Scheme III:

- Key generation:
  - A(lice) chooses a secret braid  $s \in B_n$  and computes  $b = s^2$ , so that both the CSP problem for  $s$  and  $b$ , and the RP problem for  $b$  are difficult; public key is  $b$  and private key is  $s$ .
- Authentication: Repeat the following exchanges  $k$  times:

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<sup>5</sup>In fact, Sibert et al.'s work was originally proposed in 2002 — before Cha et al.'s work published in 2003.

- A chooses a random braid  $r$ , and sends  $x = rbr^{-1}$  to B(ob);
- B sends a random bit  $\epsilon$  to A;
- For  $\epsilon = 0$ , A sends  $y = r$  to B, and B checks  $x = yby^{-1}$ ;
- For  $\epsilon = 1$ , A sends  $y = rsr^{-1}$  to B, and B checks  $x = y^2$ .

Sibert et al. proved that when the probability distribution of  $r$  picked by Alice at the authentication step is right-invariant, the above authentication schemes are zero-knowledge interactive proof of knowledge of  $s$ , under the corresponding intractability assumptions.

## 5.2.4 Braid-Based Signatures

In their pioneering paper on braid-based cryptography [175], Ko et al. pointed out that there are several challenges in finding a new digital signature scheme by using hard problems in braid groups. Two years later, Ko's team [173] reported two of the earliest braid-based signature schemes: The first scheme, denoted by SCSS, is based on the MCSP problem, and the second one, denoted by TCSS, is based on the MTSP problem. However, Ko et al. have not presented security proof on SCSS. As for TCSS, it is merely proved to be, in the random oracle model, secure against *no-message attacks*. The situation of lacking security proof for braid-based signature lasts about several years. In 2007, we [284], enlightened by the idea of “One-More-RSA-Inversion Problems” [22], defined the so-called one-more matching conjugate problem (OM-MCP) and proved that in the random oracle models both SCSS and TCSS are *existentially unforgeable against adaptively chosen message attacks (EUF-CMA)* assuming that the OM-MCP is intractable (See the next subsection).

**Protocol 5.2.5.** Simple Conjugate Signature Scheme (SCSS):

- System parameters: A noncommutative group  $G$  where CSP is intractable but CDP is tractable; A hash function  $h$  that maps a message  $m$  into an element in  $G$ .

- Key generation: A public key is a CSP-hard pair  $(x, x') \in G^2$ , and a secret key is  $a \in G$  so that  $x' = a^{-1}xa$ .
- Signing: Given a message  $m$ , a signature  $s$  is given by  $s = a^{-1}ya$  so that  $y = h(m)$ .
- Verifying: A signature  $s$  is valid if and only if  $s \sim y$  and  $x's \sim xy$ .

**Protocol 5.2.6.** Triple Conjugate Signature Scheme (TCSS):

- System parameters: A noncommutative group  $G$  where CSP is intractable but CDP is tractable; A hash function  $h$  that maps a message  $m$  into an element in  $G$ .
- Key generation: A public key is a CSP-hard pair  $(x, x') \in G^2$ , and a secret key is  $a \in G$  so that  $x' = a^{-1}xa$ .
- Signing: Given a message  $m$ , choose  $b \in G$  at random and let  $\alpha = b^{-1}xb$  and  $y = h(m||\alpha)$ , then a signature  $s$  is given by a triple  $s = (\alpha, \beta, \gamma)$  where  $\beta = b^{-1}yb$  and  $\gamma = b^{-1}aya^{-1}b$ .
- Verifying: A signature  $s = (\alpha, \beta, \gamma)$  is valid if and only if  $\alpha \sim x$ ,  $\beta \sim \gamma \sim y$ ,  $\alpha\beta \sim xy$  and  $\alpha\gamma \sim x'y$ .

### 5.2.5 One-More Like Assumptions and Provable Security

Suppose that  $N$  is the security parameter. A one-more matching conjugate problem attacker (om-mcp attacker for short) is a probabilistic polynomial-time algorithm  $\mathcal{A}$  that gets input  $p, q$  and has access to two oracles: the matching conjugate oracle  $\mathcal{O}_{mc}(\cdot)$  and the challenge oracle  $\mathcal{O}_{ch}()$ . We say that the attacker  $\mathcal{A}$  wins the game if it succeeds in matching conjugate with all  $n(N)$  braids output by the challenge oracle, but submits at most  $m(N)$  queries to the matching conjugate oracle, where  $m, n : \mathbb{N} \rightarrow \mathbb{N}$  are two polynomial functions defined over  $\mathbb{N}$  so that  $m(N) < n(N)$  holds. More formally,  $\mathcal{A}$  is invoked in the following experiment.

Experiment  $\mathbf{Exp}_{\mathbb{K}_{csp}, \mathcal{A}}^{om-mcp}(N)$

$(p, q, w) \xleftarrow{\$} \mathbb{K}_{csp}(N); n \leftarrow 0; m \leftarrow 0$

$(r_1, \dots, r_{n'}) \xleftarrow{\$} \mathcal{A}^{\mathcal{O}_{mc}, \mathcal{O}_{ch}}(p, q, N)$

If  $n' = n$  and  $m < n$  and  $\forall i = 1, \dots, n : (r_i \sim c_i) \wedge (qr_i \sim pc_i)$

Then return 1 else return 0

where the oracles are defined as

Oracle $\mathcal{O}_{mc}(b)$	Oracle $\mathcal{O}_{ch}()$
$m \leftarrow m + 1$	$n \leftarrow n + 1; c_n \xleftarrow{\$} B_N$
Return $wbw^{-1}$	Return $c_n$

The om-mcp advantage of  $\mathcal{A}$ , denoted by  $\mathbf{Adv}_{\mathbb{K}_{csp}, \mathcal{A}}^{om-mcp}(N)$ , is the probability that the above experiment returns 1, taking over the coins of  $\mathbb{K}_{csp}$ , the coins of  $\mathcal{A}$ , and the coins used by the challenge oracle across its invocations. The *one-more matching conjugate assumption* says that the one-more matching conjugate problem associated with  $\mathbb{K}_{csp}$  is hard, i.e., the function  $\mathbf{Adv}_{\mathbb{K}_{csp}, \mathcal{A}}^{om-mcp}(N)$  is negligible with respect to the security parameter  $N$  for all probabilistic polynomial-time adversaries  $\mathcal{A}$ .

Now, let us consider the modes for sampling each  $c_i$ . If the adversary  $\mathcal{A}$  can find some  $k \in \mathbb{Z}$  and  $j \in \{1, \dots, m(N) + 1\}$  so that

$$c_j = p^k. \quad (5.4)$$

Then,  $\mathcal{A}$  can set  $r_j = q^k$ . That is, the adversary can compute  $r_j$  correctly without querying the oracle  $\mathcal{O}_{mc}$  and can win the game. Similarly, if the adversary  $\mathcal{A}$  can find some  $j \in \{1, \dots, m(N) + 1\}$  and  $\alpha_i, i = 1, \dots, j - 1, j + 1, \dots, m(N) + 1$ , so that

$$c_j = \prod_{i=1, i \neq j}^{m(N)+1} c_i^{\alpha_i}, \quad (5.5)$$

then

$$r_j = \prod_{i=1, i \neq j}^{m(N)+1} r_i^{\alpha_i}. \quad (5.6)$$

Thus, the adversary can query  $c_i, i = 1, \dots, j-1, j+1, \dots, m(N)+1$ , one by one and then obtain the corresponding  $r_i$ . Now, he can also compute  $r_j$  correctly without querying the oracle again. But he can still win the game.

However, if each  $c_i$  is sampled randomly from  $B_N$ , the probability that the adversary can find  $k, j$  and  $\alpha_i, i = 1, \dots, j-1, j+1, \dots, m(N)+1$ , so that the equality (5.4) or the equality (5.5) holds is negligible with respect to the system parameter  $N$ . If  $\mathcal{A}$  tries to deduce the equation (5.5) from right to left, he has to try and try, with little advantage over guessing at random. If he tries to deduce the equation (5.5) from left to right, he has to solve a root problem, which is intractable according to [93], for each  $\alpha_i (> 0)$ . If the adversary  $\mathcal{A}$  tries to find  $k$  and  $j$  so that the equality (5.4) holds, he has to test each  $p^k, k \in \mathbb{Z}$ , by checking whether  $p^k = c_j$  holds. However, since  $\langle p \rangle$  is an infinite subgroup of  $B_N$ , the probability that the adversary can find such a pair  $(k, j)$  is also negligible with respect to  $N$ .

If the adversary is permitted to choose these  $c_i$ , he can select  $c_i$  so that the equality (5.4) or the equality (5.5) holds and this situation cannot be detected easily. Therefore, in our proposal the users, including the adversaries, are not allowed to choose these  $c_i$  by themselves. The users are allowed to perform queries on adaptively chosen messages and we employ a one-way hash function by which each message is mapped to a braid randomly sampled from  $B_N$ . The one-wayness of the hash function excludes the possibility for choosing  $c_i$  at the users' will.

**Theorem 5.2.7.** *In the random oracle model, the braid-based digital signature scheme SCSS is existentially unforgeable against adaptively chosen-message attack assuming that the one-more matching conjugate problem (OM-MCP) is intractable. More specifically, suppose that there is a forger  $\mathcal{F}$  that can  $(t, q_h, q_s, \epsilon)$ -break SCSS, then there*

exists an *om-mcp* attacker  $\mathcal{A}$  that can win the one-more matching conjugate experiment with the probability at least  $\epsilon'$  within the time  $t'$ , where

$$\epsilon' = \epsilon, \quad (5.7)$$

$$t' = t + t_s \cdot q_s + t_h \cdot q_h + t_{mc} \cdot (n(N) - q_s), \quad (5.8)$$

where  $n : \mathbb{N} \rightarrow \mathbb{N}$  is a polynomial function defined over  $\mathbb{N}$ , while  $t_s$ ,  $t_h$ , and  $t_{mc}$  are time for answering a signing oracle query, a hash oracle query and a matching conjugate oracle query, respectively.

*Proof.* That the forger  $\mathcal{F}$  can  $(t, q_h, q_s, \epsilon)$ -break SCSS means that  $\mathcal{F}$  can output a forged signature  $(m^*, r^*)$  successfully with probability at least  $\epsilon$  after he has made  $q_h$  hash queries and  $q_s$  signing queries, and then obtained the corresponding signatures  $\sigma_i = (m_i, r_i), i = 1, \dots, q_s$ . The successful forgery  $(m^*, r^*)$  means that  $r^* \sim H(m^*)$  and  $qr^* \sim pH(m^*)$  hold while  $\mathcal{F}$  has never made signing query on the message  $m^*$ . Now, let us construct another algorithm  $\mathcal{A}$  so that  $\mathcal{A}$  can win the experiment  $\mathbf{Exp}_{\text{K}_{csp}, \mathcal{A}}^{\text{om-mcp}}(N)$  with the probability at least  $\epsilon' = \epsilon$ .

Without loss of generality, we assume that  $q_h$  and  $q_s$  are bounded by  $n(N)$ . The challenge the adversary  $\mathcal{A}$  faced is, for given  $c_i, i = 1, \dots, n(N) + 1$ , to output  $r_i, i = 1, \dots, n(N) + 1$ , so that  $r_i \sim c_i$  and  $qr_i \sim pc_i$  hold, with the access to the matching conjugate oracle  $\mathcal{O}_{mc}$  at most  $n(N)$  times.

The algorithm  $\mathcal{A}$  is constructed as follows:

- **Initialize.** We define a hash list H-List as follows: H-List includes three fields *m-field*, *r-field*, and *c-field*. At the beginning, let H-List contain  $n(N) + 1$  items. For each item, the *m-field* and *r-field* are set to empty while the *c-field* is set to  $c_i$  accordingly. Then, set  $i = 0$ .
- Let  $\mathcal{A}$  play the following interactive game with the forger  $\mathcal{F}$ .

(1) Answering  $\mathcal{F}$ 's hash queries and signing queries as follows:

- Hash query on  $m$ : Locate  $m$  in  $m$ -field of H-List. If found, i.e.,  $H(m)$  has been asked before, return the corresponding  $c$ -field as the reply; otherwise, set  $i = i + 1$  and then fill  $m$  into the  $m$ -field of the  $i$ th item of the H-List, and then return the corresponding  $c$ -field as the reply.
- Signing query on  $m$ : Assume that  $\mathcal{F}$  has asked  $H(m)$  before (Otherwise, the algorithm  $\mathcal{A}$  can make such a query on behalf of the forger  $\mathcal{F}$ ). Then, there exists some  $i \in \{1, \dots, n(N) + 1\}$  so that  $c_i = H(m)$  holds. Now, let the algorithm  $\mathcal{A}$  make a query on  $c_i$  toward the matching conjugate oracle  $\mathcal{O}_{mc}$  and obtain the response  $r_i$ . Finally, let the algorithm  $\mathcal{A}$  fill  $r_i$  into the  $r$ -field of the  $i$ th item of the H-List and forward  $r_i$  as the signing reply to the forger  $\mathcal{F}$ .

Clearly, the algorithm  $\mathcal{A}$  provides perfect simulations on hash queries and signing queries for  $\mathcal{F}$ .

- (2) Suppose that after  $q_s$  times signing queries, the forger  $\mathcal{F}$  outputs a forged signature  $r^*$  on message  $m^*$ . Without loss of generality, suppose that  $\mathcal{F}$  has made a hash query on  $m^*$  (if not, let  $\mathcal{A}$  execute the query on behalf of  $\mathcal{F}$ ). Then, there exists some  $j \in \{1, \dots, n(N) + 1\}$  so that  $c_j = H(m^*)$ . Now, we set  $r_j = r^*$ .
- (3) If  $\mathcal{V}(m^*, r^*, pk) = 0$ , then let  $\mathcal{A}$  abort. Otherwise, let  $\mathcal{A}$  continue.
- (4) Locate the pair  $(m^*, r^*)$  in  $(m, r)$ -fields of H-List. If found, i.e.,  $\mathcal{F}$  has made signing query on  $m^*$  before, then let  $\mathcal{A}$  abort; otherwise, let  $\mathcal{A}$  fill  $r_j = r^*$  into the  $r$ -field of the  $j$ th item of H-List and continue.
- For each  $c_i, i = 1, \dots, j-1, j+1, \dots, n(N) + 1$ , if the  $r$ -field of the  $i$ th item of H-List is empty, let the algorithm  $\mathcal{A}$  invoke matching conjugate query  $\mathcal{O}_{mc}$  and fill the corresponding response  $r_i$  into the  $r$ -field of the  $i$ th item of H-List.



- Finally, let the algorithm  $\mathcal{A}$  read each  $r_i, i = 1, \dots, n(N) + 1$ , from the  $r$ -field of H-List and output it as the  $i$ th response for the challenge  $c_i$ . Apparently, each  $r_i$  satisfies  $(r_i \sim c_i)$  and  $(qr_i \sim pc_i)$  while  $\mathcal{A}$  has never made queries on  $c_j$  toward the matching conjugate oracle  $\mathcal{O}_{mc}$ . That is, the algorithm  $\mathcal{A}$  needs to query the matching conjugate oracle  $\mathcal{O}_{mc}$  exactly  $n(N)$  times, which is strictly less than the times that  $\mathcal{A}$  succeeds in matching conjugates with all  $n(N) + 1$  braids output by the challenge oracle. Therefore,  $\mathcal{A}$  wins the experiment  $\mathbf{Exp}_{\text{Kcsp}, \mathcal{A}}^{om-mcp}(N)$ .

The above reduction shows that as soon as the forger  $\mathcal{F}$  outputs a successful forgery (i.e., the algorithm  $\mathcal{A}$  has not aborted before the end of its running), the algorithm  $\mathcal{A}$  wins the experiment  $\mathbf{Exp}_{\text{Kcsp}, \mathcal{A}}^{om-mcp}(N)$  with probability 1. That is, the probability for  $\mathcal{A}$  to win the experiment is just the probability for  $\mathcal{F}$  to output a successful forgery, so  $\epsilon' = \epsilon$ . According to the construction of  $\mathcal{A}$ , the total running time of  $\mathcal{A}$  is  $\mathcal{F}$ 's running time plus the time for answering  $\mathcal{F}$ 's signing and hash queries and the time for  $\mathcal{A}$  to make the remained matching conjugate queries. Thus, we have  $t' = t + t_s \cdot q_s + t_h \cdot q_h + t_{mc} \cdot (n(N) - q_s)$ .  $\square$

Similarly, a provable EUF-CMA security reduction for TCSS can also be constructed by using a one-more like assumption and a random oracle model.

### 5.2.6 New Cryptographic Problems in Braid Groups

Finding new hard problems, as well as the variants, is always an interesting practice for (public-key) cryptography. In 2006, Dehornoy [94] suggested a new braid-based authentication scheme based on the shifted conjugate search problem. Let  $x, y \in B_\infty$ . The shifted conjugate operation is defined by

$$x * y = x \cdot dy \cdot \sigma_1 \cdot dx^{-1}, \quad (5.9)$$

where  $dx$  is the *shift* of  $x$  in  $B_\infty$ . That is, the operator  $d$  can be viewed as an injective function on  $B_\infty$  which sends the generator  $\sigma_i$  to the generator  $\sigma_{i+1}$  for each  $i \geq 1$ .

*Problem 5.2.8* (Shifted Conjugate Search Problem, Shifted-CSP). Let  $s, p \in B_\infty$  and  $p' = s * p$ . Find a braid  $\tilde{s}$  satisfying  $p' = \tilde{s} * p$ .

One cannot use the summit sets theory to attack Shifted-CSP, but one can apply the length-based attack [117] to try to solve Shifted-CSP. So it is interesting to study the following issues.

- (1) How to evaluate the length-based attacks on Shifted-CSP?
- (2) How to develop a theory for Shifted-CSP that will be parallel to the summit sets theory for CSP?
- (3) How to design new cryptosystems based on shifted-CSP that will be secure against all currently known attacks?

A different type of problem consists in finding the shortest words representing a given braid. We consider this problem in  $B_\infty$  which is the group generated by an infinite sequences of generators  $\{\sigma_1, \sigma_2, \dots\}$  subject to the usual braid relations.

*Problem 5.2.9* (Shortest Word Problem, SWP). Given a word  $w$  that is represented by a sequence of  $\sigma_i^{\pm 1}$  ( $i = 1, 2, \dots$ ), find the shortest word  $w'$  so that  $w' \equiv w$ .

Paterson and Razborov [226] proved that the SWP problem is co-NP-complete. This suggests us designing new cryptographic schemes in which the secret key is a short braid word, and the public key is another longer equivalent braid word. It must be noted that the co-NP-hardness of SWP holds in  $B_\infty$  only, but it is not known in  $B_n$  for fixed  $n$ . Polynomial-time algorithms for the SWP problem (with Artin presentation or Band presentation) in  $B_n$  for small fixed  $n$  were already reported [24, 25, 167, 297]. Also, an unpublished work indicates that a heuristic algorithm based on a random walk

on the Cayley graph of the braid group might give good results in solving the SWP problem [117]. Thus, the following problems are interesting.

- (4) How to design a cryptosystem based on the SWP in  $B_\infty$  by using the results from Paterson and Razborov?
- (5) How hard is the SWP problem in  $B_n$  for a fixed and moderate scale  $n$ , say  $n \geq 50$ ?

In short, to resist these attacks toward those schemes, one can try to change the distribution of the generators [112]. In this aspect, Maffre [206] proposed a new random algorithm for generating keys which are secure against the length-based attack from Hofheinz and Steinwandt [144].

## 5.3 $\mathbb{Z}$ -Modular Method

In 2006, we [73] proposed a method for building public key cryptosystems by using noncommutative rings. Given a noncommutative ring  $R$ , our proposal can work over a derived  $\mathbb{Z}$ -modular structure  $\mathbb{Z}[r]$ , or a similar  $\mathbb{Z}$ -modular structure  $\mathbb{Z}^+[r]$  (where  $r \in R$  is undetermined), provided that the corresponding derived cryptographic assumptions are reasonable. Moreover, the  $\mathbb{Z}$ -modular method is extended to noncommutative groups and noncommutative semigroups. The method developed in our schemes is called  $\mathbb{Z}$ -modular method.

Afterwards, we use  $\mathbb{Z}^+$  (resp.  $\mathbb{Z}^-$ ) to denote the set of all positive (resp. negative) integers.

### 5.3.1 $\mathbb{Z}$ -Modular Method over Noncommutative Rings

In this subsection, we give a very simple review on basic concepts about  $\mathbb{Z}$ -modular over noncommutative rings. Readers who know these concepts can go to the next sub-

section directly.

Suppose that  $R$  is a ring with  $(R, +, \mathbf{0})$  and  $(R, \cdot, \mathbf{1})$  as its additive abelian group and multiple non-abelian semigroup, respectively. Let us consider integral coefficient polynomials with ring assignment. The notion of scalar multiplication over  $R$  is already well-defined. For  $k \in \mathbb{Z}^+$  and  $r \in R$ ,

$$(k)r \triangleq \underbrace{r + \cdots + r}_{k \text{ times}}. \quad (5.10)$$

When  $k \in \mathbb{Z}^-$ , we define

$$(k)r \triangleq (-k)(-r) = \underbrace{(-r) + \cdots + (-r)}_{-k \text{ times}}. \quad (5.11)$$

For  $k = 0$ , it is natural to define  $(k)r = \mathbf{0}$ .

**Proposition 5.3.1.**  $(a)r^m \cdot (b)r^n = (ab)r^{m+n} = (b)r^n \cdot (a)r^m, \forall a, b, m, n \in \mathbb{Z}$  and  $\forall r \in R$ .

*Proof.* According to the definition of scalar multiplication, the distributivity of multiplication with respect to addition, and commutativity of addition, this statement is concluded immediately.  $\square$

*Remark 5.3.2.* In general,  $(a)r \cdot (b)s \neq (b)s \cdot (a)r$  when  $r \neq s$ , since multiplication in  $R$  is noncommutative.

Recall a polynomial with positive integral coefficient  $f(x) = a_0 + a_1x + \cdots + a_nx^n \in \mathbb{Z}^+[x]$ . If we assign the undetermined element  $x$  as an element  $r \in R$ , then we obtain a well-defined element in  $R$  as shown below:

$$f(r) = \sum_{i=0}^n (a_i)r^i = (a_0)\mathbf{1} + (a_1)r + \cdots + (a_n)r^n. \quad (5.12)$$

Further, let  $r$  be undetermined, then  $f(r)$  is univariable polynomial over  $R$ . The set of all (positive integral coefficient) univariable polynomials over  $R$  is denoted by  $\mathbb{Z}^+[r]$ .

Suppose that  $f(r) = \sum_{i=0}^n (a_i)r^i \in \mathbb{Z}^+[r]$ ,  $h(r) = \sum_{j=0}^m (b_j)r^j \in \mathbb{Z}^+[r]$  and  $n \geq m$ , then

$$\left( \sum_{i=0}^n (a_i)r^i \right) + \left( \sum_{j=0}^m (b_j)r^j \right) = \left( \sum_{i=0}^m (a_i + b_i)r^i \right) + \left( \sum_{i=m+1}^n (a_i)r^i \right), \quad (5.13)$$

and according to Property 5.3.1 as well as the distributivity, we have

$$\left( \sum_{i=0}^n (a_i)r^i \right) \cdot \left( \sum_{j=0}^m (b_j)r^j \right) = \sum_{i=0}^{n+m} (p_i)r^i, \quad (5.14)$$

where  $p_i = \sum_{j=0}^i a_j b_{i-j} = \sum_{j+k=i} a_j b_k$ . And then, we have the following theorem.

**Theorem 5.3.3.**  $f(r) \cdot h(r) = h(r) \cdot f(r), \forall f(r), h(r) \in \mathbb{Z}^+[r]$ .

### 5.3.2 New Problems over Noncommutative Rings

We would like to introduce the following cryptographic problems over a noncommutative group  $G$ :

- **Symmetrical Decomposition Problem (SDP):** Given  $(x, y) \in G \times G$  and  $m, n \in \mathbb{Z}$ , find  $z \in G$  so that  $y = z^m x z^n$ .
- **Generalized Symmetrical Decomposition Problem (GSDP):** Given  $(x, y) \in G \times G$ ,  $S \subseteq G$  and  $m, n \in \mathbb{Z}$ , find  $z \in S$  so that  $y = z^m x z^n$ .

Clearly, GSDP can be viewed as a constrained variation of SDP. In general, if the size of  $S$  is large enough and its membership information *does not* help one to extract  $z$  from  $z^m x z^n$ , we believe that GSDP is at least as hard as SDP.

In the above definition of GSDP, if the parameters  $m$  and  $n$  are fixed, we define a new function  $\hat{e}_{m,n}$  by

$$\hat{e}_{m,n} : G \times S, \quad (x, z) \mapsto z^m x z^n. \quad (5.15)$$

Further, if we denote  $z^m x z^n$  as a new form  $x^z$ , then the above function can be viewed as a newly introduced exponential operation on  $G$  with respect to its subset  $S$ <sup>6</sup>. Similarly, if  $y = z^m x z^n$ , then  $z$  can be viewed as the discrete logarithm of  $y$  with respect to the base  $x$ , i.e.,  $z$  can be denoted by  $\log_x y$ .

Now, we can regard GSDP as the discrete logarithm (DL) problem over  $G$ . Based on this observation, we introduce a new generalized computational Diffie–Hellman (GCDH) problem over  $G$ :

- **Generalized Computational Diffie–Hellman Problem (GCDH):** Compute  $x^{z_1 z_2}$  (or  $x^{z_2 z_1}$ ) given  $x, x^{z_1}, x^{z_2} \in G$ , where  $G$  is a noncommutative group and  $z_1, z_2 \in S \subseteq G$ .

Note that if  $z_1$  lies in the center of  $z_2$ , i.e.,  $z_1$  commutes  $z_2$ ,  $x^{z_1 z_2} = x^{z_2 z_1}$  holds. It is clear that if GSDP, i.e., DL problem over  $G$  is tractable, so is GCDH problem over  $G$ . But the converse is not true. At present, there is no clue to solve this kind of GCDH problem without extracting  $z_1$  (or  $z_2$ ) from  $x$  and  $x^{z_1}$  (or  $x^{z_2}$ ).

The GCDH assumption over  $G$  says that GCDH problem over  $G$  is intractable, i.e., there is no probabilistic polynomial time algorithm which can solve GCDH problem over  $G$  with non-negligible advantage with respect to the parameters of the problem scale.

Likewise, we arrive at the concept of the GSD (also DL) and GCDH assumptions over a noncommutative semigroup  $G$  and  $m, n \in \mathbb{Z}^-$ .

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<sup>6</sup>In sequel we omit the clause of “with respect to its subset  $S$ ” for visual comfort, unless the set  $S$  is explicitly specified.

Now, suppose that  $(R, +, \cdot)$  is a noncommutative ring. For any  $a \in R$ , we define a set  $P_a \subseteq R$  by

$$P_a \triangleq \{f(a) : f(r) \in \mathbb{Z}^+[r]\}.$$

Then, let us consider the new versions of GSD and GCDH problems over  $(R, \cdot)$  with respect to its subset  $P_a$ , and name them as polynomial symmetric decomposition (PSD) problem and polynomial Diffie–Hellman (PDG) problem respectively:

- **Polynomial Symmetrical Decomposition (PSD) Problem:** Given  $(a, x, y) \in R^3$  and  $m, n \in \mathbb{Z}$ , find  $z \in P_a$  so that  $y = z^m x z^n$ , where  $R$  is a noncommutative ring.
- **Polynomial Diffie–Hellman (PDH) Problem:** Compute  $x^{z_1 z_2}$  (or  $x^{z_2 z_1}$ ) for given  $a, x, x^{z_1}$ , and  $x^{z_2}$ , where  $a, x \in R, z_1, z_2 \in P_a$ , and  $R$  is a noncommutative ring.

Accordingly, the PSD (PDH, respectively) cryptographic assumption says that PSD (PDH, respectively) problem over the ring  $(R, \cdot)$  is intractable, i.e., there does not exist probabilistic polynomial time algorithm which can solve PSD (PDH, respectively) problem over  $(R, \cdot)$  with non-negligible advantage with respect to the parameters of the problem scale.

### 5.3.3 Diffie–Hellman-Like Key Agreement Protocol

Now, let us consider Diffie–Hellman-like key agreement protocol over a noncommutative ring  $R$ .

**Protocol 5.3.4.** Diffie–Hellman-Like Key Agreement Protocol over Noncommutative Ring:

- (0) Alice sends two random small, positive integers (say, less than 10)  $m, n \in \mathbb{Z}^+$  and two random elements  $a, b \in R$  to Bob.

- (1) Alice chooses a random polynomial  $f(x) \in \mathbb{Z}^+[x]$  so that  $f(a) \neq \mathbf{0}$  and then takes  $f(a)$  as her private key.
- (2) Bob chooses a random polynomial  $h(x) \in \mathbb{Z}^+[x]$  so that  $h(a) \neq \mathbf{0}$  and then takes  $h(a)$  as his private key.
- (3) Alice computes  $r_A = f(a)^m \cdot b \cdot f(a)^n$  and sends  $r_A$  to Bob.
- (4) Bob computes  $r_B = h(a)^m \cdot b \cdot h(a)^n$  and sends  $r_B$  to Alice.
- (5) Alice computes  $K_A = f(a)^m \cdot r_B \cdot f(a)^n$  as the shared session key.
- (6) Bob computes  $K_B = h(a)^m \cdot r_A \cdot h(a)^n$  as the shared session key.

In practice, the steps (0), (1), and (3) can be finished simultaneously and require only one pass communication from Alice to Bob. After that, the steps (2) and (4) can be finished in one pass communication from Bob to Alice. Finally, Alice and Bob can execute the steps (5) and (6), respectively. See Figure 5.4 for more details.

Pass	Alice	Bob
	$m, n \xleftarrow{\$} \mathbb{Z}^+$ $a, b \xleftarrow{\$} R$ $f(x) \xleftarrow{\$} \mathbb{Z}^+[x]$	
1	$\xrightarrow{m, n, a, b, f(a)^m b f(a)^n}$	
	$\xleftarrow{h(x) \xleftarrow{\$} \mathbb{Z}^+[x]}$	
2	$\xleftarrow{h(a)^m b h(a)^n}$	
	$K_A = f(a)^m h(a)^m b h(a)^n f(a)^n = K_B = h(a)^m f(a)^m b f(a)^n h(a)^n$	

Figure 5.4: Diffie–Hellman-Like Key Agreement over Noncommutative Rings

Clearly, the above key agreement protocol can resist passive adversary under the PDH assumption over the noncommutative monoid  $(R, \cdot)$ . It is similar to the standard Diffie–Hellman protocol that the protocol depicted in Figure 5.4 cannot resist the man-in-the-middle (MIM) attack. But it is easy to improve it so that the improvement can resist the MIM attack. This is left for interested readers.



### 5.3.4 ElGamal-Like Encryption Scheme

By the above key agreement protocol, we have the following ElGamal-like encryption scheme.

**Protocol 5.3.5.** ElGamal-Like Encryption (Basic) Scheme over Noncommutative Ring:

- **Setup:** Suppose that SDP is intractable on the monoid  $(R, \cdot)$ , where  $(R, +, \cdot)$  is a noncommutative ring. Pick two small positive integers  $m, n \in \mathbb{Z}^+$ . Let  $H : R \rightarrow \mathcal{M}$  be a cryptographic hash function which maps  $R$  to the message space  $\mathcal{M}$ . Set the public parameters of the system as the tuple  $\langle R, m, n, \mathcal{M}, H \rangle$ .
- **Key generation:** Each user chooses two random elements  $p, q \in R$  and a random polynomial  $f(x) \in \mathbb{Z}^+[x]$  so that  $f(p) \neq \mathbf{0}$  and then takes  $f(p)$  as his private key, computes  $y = f(p)^m \cdot q \cdot f(p)^n$  and publishes his public key  $(p, q, y) \in R^3$ .
- **Encryption:** Given a message  $M \in \mathcal{M}$  and the receiver's key  $(p, q, y) \in R^3$ , the sender chooses a random polynomial  $h(x) \in \mathbb{Z}^+[x]$  so that  $h(p) \neq \mathbf{0}$  and computes

$$c = h(p)^m \cdot q \cdot h(p)^n, \quad d = H(h(p)^m \cdot y \cdot h(p)^n) \oplus M.$$

Output the ciphertext  $(c, d) \in R \times \mathcal{M}$ .

- **Decryption:** Upon receiving a ciphertext  $(c, d) \in R \times \mathcal{M}$ , the receiver, by using his private key  $f(p)$ , computes the plaintext

$$M = H(f(p)^m \cdot c \cdot f(p)^n) \oplus d$$

Now, we prove that the above basic encryption scheme is “all-or-nothing” secure.

The proof is very similar to that of Theorem 8.3 [209].

**Theorem 5.3.6.** *For a plaintext message uniformly distributed in the plaintext message space, the above encryption is “all-or-nothing” secure against CPA under the PDH assumption over the noncommutative ring  $(R, +, \cdot)$  provided that  $H$  is a random oracle.*

*Proof.* On one hand, if PDH problem is tractable for any given ciphertext pair  $(c, d)$  and the corresponding public key  $(p, q, y)$ , it is easy to extract  $k = q^{(\log_q c)(\log_q y)}$  from the triple  $(q, c, y)$  and then compute the plaintext  $M = d \oplus H(k)$ .

On the other hand, suppose that there exists an efficient adversary  $\mathcal{A}$ , with access to the random oracle  $H$ , who is able to break the above cryptosystem, i.e., given any public key  $(p, q, y = f(p)^m q f(p)^n)$  and ciphertext  $(c, d)$ ,  $\mathcal{A}$  outputs

$$M \leftarrow \mathcal{A}^H(p, q, y, c, d)$$

with a non-negligible advantage  $\epsilon$  so that  $M$  satisfies

$$M = d \oplus H(y^{\log_q c}) = d \oplus H(q^{(\log_q y)(\log_q c)}),$$

i.e.,

$$M = d \oplus H(h^m y h^n) \text{ and } c = h^m q h^n$$

for some  $h \in P_p$ . Then, for an arbitrary PDH instance  $(a, x, x^{z_1}, x^{z_2})$ , we set  $(a, x, x^{z_1})$  as public key and  $(x^{z_2}, d)$  as ciphertext pair for a random  $d \in \mathcal{M}$ . Then, with the advantage  $\epsilon$ ,  $\mathcal{A}$  outputs

$$M \leftarrow \mathcal{A}^H(a, x, x^{z_1}, x^{z_2}, d)$$

so that

$$M = d \oplus H(x^{z_1 z_2}), \text{ i.e., } M = d \oplus H(z_2^m z_1^m x z_1^n z_2^n)$$

for some  $z_2 \in P_a$ . Recall that  $z_1 \in P_a$ , thus  $z_2 z_1 = z_1 z_2$  by Theorem 5.3.3. Then,

$$M = d \oplus H(z_2^m z_1^m x z_1^n z_2^n) = d \oplus H(x^{z_1 z_2}) = d \oplus H(x^{z_2 z_1}).$$

Clearly, if the adversary  $\mathcal{A}$ 's advantage  $\epsilon$  is non-negligible, then  $\mathcal{A}$  must make corresponding  $H$ -query on  $x^{z_1 z_2}$ . Otherwise, without knowing the hash value  $H(x^{z_1 z_2})$ ,  $\mathcal{A}$ 's advantage for computing correct  $M$  should be *negligible*, since  $H$  is modeled as a cryptographic hash.

With the random oracle assumption on  $H$ , we can setup a  $H$ -list which contains two fields  $(r_i, h_i)$  and is initialized with empty. Whenever the adversary  $\mathcal{A}$  makes a  $H$ -query with input  $r$ , we examine whether there exists the pair  $(r, h)$  in  $H$ -list. If so, return  $h$  as the answer to  $\mathcal{A}$ ; otherwise, randomly pick  $h \in \mathcal{M}$ , put the pair  $(r, h)$  into  $H$ -list and return  $h$  as the answer to  $\mathcal{A}$ . Clearly, the simulation on  $H$  is perfect. Finally, when  $\mathcal{A}$  outputs  $M$ , we can retrieve the correct item  $x^{z_1 z_2} = r_i$  by checking the equality  $M = d \oplus h_i$ . Thus, we can solve PDH problem with the non-negligible probability  $\epsilon$ . This contradicts the PDH assumption.  $\square$

The above theorem shows that the basic scheme is of the weakest security, i.e., the OW-CPA security. Although we can use a technique proposed by Fujisaki and Okamoto (at CRYPTO'99) [116] to convert the above basic scheme into a chosen ciphertext secure system in the random oracle, we would like to adopt another technique also proposed by Fujisaki and Okamoto (at PKC'99) [115] because the latter is more compact.

First, we need to prove that the above basic scheme is of IND-CPA security.

**Theorem 5.3.7.** *Let  $H$  be a random oracle from  $R$  to  $\mathcal{M}$ . Let  $\mathcal{A}$  be an IND-CPA adversary that has advantage  $\epsilon$  against the above basic scheme within  $t$  steps. Suppose  $\mathcal{A}$  makes a total of  $q_H > 0$  queries to  $H$ . Then there is an algorithm  $\mathcal{B}$  that solves PDH problem over the noncommutative ring  $R$  with advantage at least  $\epsilon'$  within  $t'$  steps,*

where

$$\epsilon' = \frac{2\epsilon}{q_H}, \text{ and } t' = \mathcal{O}(t).$$

*Proof.* The algorithm  $\mathcal{B}$  takes as input a 4-tuple  $(a, x, y_1, y_2)$  with  $y_i = x^{z_i} = z_i^m x z_i^n$  for unknown  $z_i \in P_a$ ,  $i = 1, 2$ , i.e., a PDH instance. Let  $y = x^{z_1 z_2}$  denote the solution to the PDH instance.

- **Setup.**  $\mathcal{B}$  sets the system parameters as  $\langle R, m, n, \mathcal{M}, H \rangle$  and  $(a, x, y_1)$  as a public key.
- **H-queries.**  $\mathcal{B}$  sets up a  $H$ -list which contains two fields  $(r_j, h_j)$  and is initialized with empty. Whenever the adversary  $\mathcal{A}$  makes a  $H$ -query with input  $r$ ,  $\mathcal{B}$  examines whether there exists the pair  $(r, h)$  in  $H$ -list. If so, return  $h$  as the answer to  $\mathcal{A}$ ; otherwise, randomly pick  $h \in \mathcal{M}$ , put the pair  $(r, h)$  into  $H$ -list and return  $h$  as the answer to  $\mathcal{A}$ . Clearly, the simulation on  $H$  is perfect.
- **Challenge.** When  $\mathcal{A}$  outputs two messages  $M_0$  and  $M_1$ ,  $\mathcal{B}$  picks randomly a string  $d \in \mathcal{M}$  and sets  $C$  as the ciphertext pair  $(y_2, d)$ . It then gives  $C$  to  $\mathcal{A}$  as the challenge. Notice that, the plaintext corresponding to  $C$  is  $d \oplus H(x^{(\log_x y_1)(\log_x y_2)}) = d \oplus H(x^{z_1 z_2}) = d \oplus H(y)$ . (Recall that  $z_1, z_2$ , and  $y$  are all unknown to  $\mathcal{A}$  and  $y$  is just the solution to the above PDH instance.)
- **Guess.**  $\mathcal{A}$  outputs its guess  $b' \in \{0, 1\}$ . At this point,  $\mathcal{B}$  picks a random tuple  $(r_j, h_j)$  from the  $H$ -list and outputs  $r_j$  as the solution to the given PDH instance.

It is easy to see that  $\mathcal{A}$ 's view in  $\mathcal{B}$ 's simulation is the same as that in a real attack, in other words, the simulation is perfect. So  $\mathcal{A}$ 's advantage in this simulation will be  $\epsilon$ .

We let  $\mathcal{H}$  be the event that  $y$  is queried to  $H$  oracle during  $\mathcal{B}$ 's simulation. Notice that  $H(y)$  is independent of  $\mathcal{A}$ 's view. If  $\mathcal{A}$  never queries  $y$  to the  $H$  oracle in the above simulation, the plaintext corresponding to  $C$  is also independent of  $\mathcal{A}$ 's view. Therefore, in the simulation we have  $\Pr[b = b' | \neg \mathcal{H}] = 1/2$ . We know that in the real attack (and

also in the simulation)  $|\Pr[b = b'] - 1/2| \geq \epsilon$ . We have

$$\begin{aligned}
 \Pr[b = b'] &= \Pr[b = b' | \neg \mathcal{H}] \Pr[\neg \mathcal{H}] + \Pr[b = b' | \mathcal{H}] \Pr[\mathcal{H}] \\
 &\leq \Pr[b = b' | \neg \mathcal{H}] \Pr[\neg \mathcal{H}] + \Pr[\mathcal{H}] \\
 &= \frac{1}{2} \Pr[\neg \mathcal{H}] + \Pr[\mathcal{H}] \\
 &= \frac{1}{2} + \frac{1}{2} \Pr[\mathcal{H}],
 \end{aligned}$$

$$\begin{aligned}
 \Pr[b = b'] &\geq \Pr[b = b' | \neg \mathcal{H}] \Pr[\neg \mathcal{H}] \\
 &= \frac{1}{2} \Pr[\neg \mathcal{H}] \\
 &= \frac{1}{2} (1 - \Pr[\mathcal{H}]) \\
 &= \frac{1}{2} - \frac{1}{2} \Pr[\mathcal{H}].
 \end{aligned}$$

Hence,  $|\Pr[b = b'] - 1/2| \leq \frac{1}{2} \Pr[\mathcal{H}]$ . Since  $|\Pr[b = b'] - 1/2| \geq \epsilon$ ,  $\Pr[\mathcal{H}] \geq 2\epsilon$ . Furthermore, by the definition of the event  $\mathcal{H}$ , we know that  $y$  is in some tuple on the  $H$ -list with probability at least  $2\epsilon$ . It follows that  $\mathcal{B}$  outputs the correct answer to the above PDH instance with probability at least  $2\epsilon/q_H$ .  $\square$

At PKC'99, Fujisaki and Okamoto [115] introduced a method to convert an IND-CPA encryption scheme into an IND-CCA2 scheme. For self-contained, we rephrase their main idea as follows.

Suppose that  $\Pi := \{\mathcal{K}, \mathcal{E}, \mathcal{D}\}$  is an IND-CPA secure public-key encryption scheme with key generation algorithm  $\mathcal{K}(1^k)$ , encryption algorithm  $\mathcal{E}_{pk}(m, s)$  and decryption algorithm  $\mathcal{D}_{sk}(y)$ , where  $pk$  and  $sk$  are a public key and the corresponding private key,  $m$  is a message with  $k + k_0$  bits,  $s$  is a random string with  $l$  bits and  $y$  is a ciphertext.

The converted public-key encryption scheme  $\bar{\Pi} := \{\bar{\mathcal{K}}, \bar{\mathcal{E}}, \bar{\mathcal{D}}\}$  is defined by

$$\begin{aligned}\bar{\mathcal{K}}(1^k) &:= \mathcal{K}(1^{k+k_0}), \\ \bar{\mathcal{E}}_{pk}(x, r) &:= \mathcal{E}_{pk}(x \parallel r, H(x \parallel r)), \\ \bar{\mathcal{D}}_{sk}(y) &:= \begin{cases} MSB_k(\mathcal{D}_{sk}(y)), & \text{if } y = \mathcal{E}_{pk}(\mathcal{D}_{sk}(y), H(\mathcal{D}_{sk}(y))) \\ \perp, & \text{otherwise} \end{cases}\end{aligned}$$

where  $MSB_k(\cdot)$  returns the leading  $k$  bits of the input bit-string,  $H$  is a random function of  $\{0, 1\}^{k+k_0} \rightarrow \{0, 1\}^l$ ,  $x$  is a message with  $k$  bits,  $r$  is a random string with  $k_0$  bits and  $\parallel$  denotes concatenation.

**Theorem 5.3.8** (Fujisaki-Okamoto Theorem [115]). *Suppose that  $\Pi(1^{k+k_0})$  is an IND-CPA secure scheme and  $\bar{\Pi}$  is the converted scheme. If there exists a  $(t, q_H, q_D, \epsilon)$ -breaker  $\mathcal{A}$  for  $\bar{\Pi}(1^k)$  in the sense of IND-CCA2 in the random oracle model, there exist constant  $c$  and a  $(t', 0, 0, \epsilon')$ -breaker  $\mathcal{A}'$  for  $\Pi(1^{k+k_0})$  where*

$$\begin{aligned}\epsilon' &= (\epsilon - q_H \cdot 2^{-(k_0-1)}) \cdot (1 - 2^{-l_0})^{q_D} \text{ and} \\ t' &= t + q_H \cdot (T_{\mathcal{E}}(k) + c \cdot k).\end{aligned}$$

Here,  $(t, q_H, q_D, \epsilon)$ -breaker  $\mathcal{A}$ , means that  $\mathcal{A}$  stops within  $t$  steps, succeeds with probability at least  $\epsilon$ , makes at most  $q_H$  queries to random oracle  $H$ , and makes at most  $q_D$  queries to decryption oracle  $D_{sk}$ .  $T_{\mathcal{E}}(k)$  denotes the computational time of the encryption algorithm  $\mathcal{E}_{pk}(\cdot)$ , and

$$l_0 := \log_2\left(\min_{x \in \{0,1\}^{k+k_0}} \{\#\{E_{pk}(x, r) \mid r \in \{0,1\}^l\}\}\right).$$

*Proof.* See Theorem 3 in [115]. □

According to Fujisaki-Okamoto [115], we can convert our basic encryption into a

new which is of IND-CCA2 security by padding each plaintext with  $k_0$  random bits. The new scheme has the same system parameters  $\langle R, m, n, \mathcal{M} \rangle$ . The cryptographic hash function  $H$  in basic scheme is replaced with two new cryptographic hash functions  $H_1 : \{0, 1\}^{k+k_0} \rightarrow \mathbb{Z}^+[x]$  and  $H_2 : R \rightarrow \{0, 1\}^{k+k_0}$ , where  $k$  is the length of an original message.

**Protocol 5.3.9.** Elgamal-Like Encryption (Improved) Scheme over Non-commutative Ring:

- **Setup:** System public parameters include  $\langle R, m, n, \mathcal{M} \rangle$  and  $k_0, H_1, H_2$ .
- **Key generation:** See the basic scheme.
- **Encryption:** Given a message  $M \in \mathcal{M}$  and receiver's key  $(p, q, y = f(p)^m \cdot q \cdot f(p)^n) \in R^3$ , the sender chooses a random  $r \in \{0, 1\}^{k_0}$  and *extracts*<sup>7</sup> a polynomial  $h(x) = H_1(M \parallel r) \in \mathbb{Z}^+[x]$  so that  $h(p) \neq \mathbf{0}$  and then computes

$$c = h(p)^m \cdot q \cdot h(p)^n, \quad d = H_2(h(p)^m \cdot pk \cdot h(p)^n) \oplus (M \parallel r).$$

Output the ciphertext  $(c, d) \in R \times \{0, 1\}^{k+k_0}$ .

- **Decryption:** Upon receiving a ciphertext  $(c, d) \in R \times \{0, 1\}^{k+k_0}$ , the receiver, by using his private key  $f(p)$ , computes

$$M' = H_2(f(p)^m \cdot c \cdot f(p)^n) \oplus d.$$

Extract  $g(x) = H_1(M') \in \mathbb{Z}^+[x]$  and check whether  $c = g(p)^m \cdot q \cdot g(p)^n$  holds.

If so, output the first  $k$  bits of  $M'$ ; otherwise, output  $\perp$ .

By Theorem 5.3.7 and Theorem 5.3.8, we have

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<sup>7</sup>See Remark 5.3.5 for further discussion.

**Theorem 5.3.10.** *Let  $H_1$  and  $H_2$  be random oracles. Then the new scheme is an adaptively chosen ciphertext secure (IND-CCA2) encryption under the PDH assumption. More specifically, suppose there is an IND-CCA2 adversary  $\mathcal{A}$  that has advantage  $\epsilon$  against the new scheme within  $t$  steps. Suppose  $\mathcal{A}$  makes at most  $q_D$  decryption queries, and at most  $q_{H_1}$ ,  $q_{H_2}$  queries to the hash functions  $H_1$ ,  $H_2$ , respectively. Then there is an algorithm  $\mathcal{B}$  which can solve PDH with the probability at least  $\epsilon'$  within  $t'$  steps, where*

$$\begin{aligned}\epsilon' &= \frac{2}{q_{H_1}} \left\{ \frac{\epsilon}{(1 - 2^{-l_0})^{q_D}} + q_{H_2} \cdot 2^{-(k_0-1)} \right\}, \text{ and} \\ t' &= \mathcal{O}(t - q_{H_2} \cdot (T_{\mathcal{E}}(k) + c \cdot k))\end{aligned}$$

where  $c$  is a constant and  $T_{\mathcal{E}}(k)$  denotes the computational time of the encryption algorithm  $\mathcal{E}_{pk}(\cdot)$  in our basic scheme, and

$$l_0 := \log_2 \left( \min_{x \in \{0,1\}^{k+k_0}} [\#\{E_{pk}(x, r) | r \in \{0,1\}^l\}] \right).$$

*Proof.* By the theorems 5.3.7 and 5.3.8, it can be immediately concluded that the new encryption is of IND-CCA2 security in the random oracle model under the PDH assumption.  $\square$

### 5.3.5 Instantiation and Illustration (I)

In this subsection, we describe the method to construct a cryptographic hash that maps a binary string to a polynomial, such as  $H_1 : \{0, 1\}^{k+k_0} \rightarrow \mathbb{Z}^+[x]$ . In particular, the resulting polynomials should satisfy more conditions, such as the condition  $h(p) \neq \mathbf{0}$  and so on. We employ the so-called *divide-and-conquer* strategy to solve this problem: At first, we extract a polynomial  $h(x) \in \mathbb{Z}^+[x]$  from a binary string in  $\{0, 1\}^{k+k_0}$ ; Then, we adopt a deterministic way to transform  $h(x)$  to  $\tilde{h}(x)$  so that  $\tilde{h}(x)$  satisfies the desired



condition  $C$ . In other words, we need to consider the following issues:

- **Initialization.** Suppose that there is a cryptographic hash function  $H_1$  which maps  $m \in \{0, 1\}^{k+k_0}$  to  $(z_0, z_1, \dots, z_{d_H}) \in \mathbb{Z}_{c_M}^{d_H+1}$ , where  $d_H \cdot c_M$  is large enough to resist brute force attack.
- **Transformation.** For a polynomial  $h(x) = z_0 + z_1x + \dots + z_{d_H}x^{d_H}$ , it can be transformed to  $\tilde{h}(x) = h(x) + \Delta$  with

$$\Delta = \min\{\delta \in \mathbb{Z}_{\geq 0} : h(x) + \delta \cdot \mathbf{1}_R \in \mathbb{Z}^+[x] \cap C\},$$

where  $\mathbb{Z}^+[x] \cap C$  is the set of polynomials in  $\mathbb{Z}^+[x]$  satisfying the given condition  $C$ .

Now, let us illustrate the above method by using a special matrix ring:  $M_2(\mathbb{Z}_N)$ , where  $N = p \cdot q$ ,  $p$  and  $q$  are two large secure primes. We have solid reasons to believe that SDP over  $M_2(\mathbb{Z}_N)$  is intractable, since it is infeasible to extract

$$A = \begin{pmatrix} a & 0 \\ 0 & 0 \end{pmatrix} \in M_2(\mathbb{Z}_N), a \in \mathbb{Z}_N$$

from

$$A^2 = \begin{pmatrix} a^2 \bmod N & 0 \\ 0 & 0 \end{pmatrix} \in M_2(\mathbb{Z}_N)$$

without knowing the factoring of  $N$ .

**Example 1.** Diffie–Hellman–Like key agreement over matrix rings. Let  $N = 7 \cdot 11$ . Alice chooses

$$m = 3, n = 5, A = \begin{pmatrix} 2 & 5 \\ 7 & 4 \end{pmatrix}, B = \begin{pmatrix} 1 & 9 \\ 3 & 2 \end{pmatrix}, \text{ and } f(x) = 3x^3 + 4x^2 + 5x + 6.$$

She computes

$$f(A) = 3 \cdot \begin{pmatrix} 2 & 5 \\ 7 & 4 \end{pmatrix}^3 + 4 \cdot \begin{pmatrix} 2 & 5 \\ 7 & 4 \end{pmatrix}^2 + 5 \cdot \begin{pmatrix} 2 & 5 \\ 7 & 4 \end{pmatrix} + 6 \cdot I = \begin{pmatrix} 35 & 12 \\ 63 & 9 \end{pmatrix},$$

and

$$r_A = \begin{pmatrix} 35 & 12 \\ 63 & 9 \end{pmatrix}^3 \begin{pmatrix} 1 & 9 \\ 3 & 2 \end{pmatrix} \begin{pmatrix} 35 & 12 \\ 63 & 9 \end{pmatrix}^5 = \begin{pmatrix} 49 & 53 \\ 42 & 31 \end{pmatrix}.$$

Then, she sends  $m$ ,  $n$ ,  $A$ ,  $B$ , and  $r_A$  to Bob. Upon receiving  $m$ ,  $n$ ,  $A$ ,  $B$ , and  $r_A$  from Alice, Bob chooses another polynomial  $h(x) = x^5 + 5x + 1$  and computes

$$h(A) = \begin{pmatrix} 2 & 5 \\ 7 & 4 \end{pmatrix}^5 + 5 \cdot \begin{pmatrix} 2 & 5 \\ 7 & 4 \end{pmatrix} + I = \begin{pmatrix} 64 & 13 \\ 49 & 23 \end{pmatrix},$$

and

$$r_B = \begin{pmatrix} 64 & 13 \\ 49 & 23 \end{pmatrix}^3 \begin{pmatrix} 1 & 9 \\ 3 & 2 \end{pmatrix} \begin{pmatrix} 64 & 13 \\ 49 & 23 \end{pmatrix}^5 = \begin{pmatrix} 29 & 40 \\ 52 & 6 \end{pmatrix}.$$

Then, he sends  $r_B$  to Alice. Finally, Alice extracts the session key as

$$K_A = \begin{pmatrix} 35 & 12 \\ 63 & 9 \end{pmatrix}^3 \begin{pmatrix} 29 & 40 \\ 52 & 6 \end{pmatrix} \begin{pmatrix} 35 & 12 \\ 63 & 9 \end{pmatrix}^5 = \begin{pmatrix} 28 & 37 \\ 14 & 40 \end{pmatrix},$$

and Bob extracts the session key as

$$K_B = \begin{pmatrix} 64 & 13 \\ 49 & 23 \end{pmatrix}^3 \begin{pmatrix} 49 & 53 \\ 42 & 31 \end{pmatrix} \begin{pmatrix} 64 & 13 \\ 49 & 23 \end{pmatrix}^5 = \begin{pmatrix} 28 & 37 \\ 14 & 40 \end{pmatrix}.$$

**Example 2.** Encryption/Decryption over matrix rings for the basic scheme. Suppose that in the basic scheme, the message space is  $M_2(\mathbb{Z}_N)$  and

$$H : M_2(\mathbb{Z}_N) \rightarrow M_2(\mathbb{Z}_N), m_{ij} \mapsto 2^{m_{ij}} \bmod N,$$

is a cryptographic hash function. Suppose that the remained system parameters are

$$m = 3, n = 5, p = \begin{pmatrix} 2 & 5 \\ 7 & 4 \end{pmatrix}, q = \begin{pmatrix} 1 & 9 \\ 3 & 2 \end{pmatrix}.$$

Suppose that the polynomial  $f(x)$  picked by Alice is just that of in example 1. That is, Alice's private key is  $f(p) = \begin{pmatrix} 35 & 12 \\ 63 & 9 \end{pmatrix}$ , and the corresponding public key is  $pk \triangleq f(p)^3 q f(p)^5 = \begin{pmatrix} 49 & 53 \\ 42 & 31 \end{pmatrix}$ . Bob picks a random message  $M = \begin{pmatrix} 27 & 19 \\ 34 & 8 \end{pmatrix}$ . Suppose the random polynomial he picked is  $h(p) = \begin{pmatrix} 64 & 13 \\ 49 & 23 \end{pmatrix}$ . Now, he computes the ciphertext  $(c, d)$  as follows:

$$c = h(p)^3 q h(p)^5 = \begin{pmatrix} 64 & 13 \\ 49 & 23 \end{pmatrix}^3 \begin{pmatrix} 1 & 9 \\ 3 & 2 \end{pmatrix} \begin{pmatrix} 64 & 13 \\ 49 & 23 \end{pmatrix}^5 = \begin{pmatrix} 29 & 40 \\ 52 & 6 \end{pmatrix},$$

$$\begin{aligned} d &= H(h(p)^3 \cdot pk \cdot h(p)^5) \oplus M \\ &= H\left(\begin{pmatrix} 64 & 13 \\ 49 & 23 \end{pmatrix}^3 \begin{pmatrix} 49 & 53 \\ 42 & 31 \end{pmatrix} \begin{pmatrix} 64 & 13 \\ 49 & 23 \end{pmatrix}^5\right) \oplus \begin{pmatrix} 27 & 19 \\ 34 & 8 \end{pmatrix} \\ &= \left(\begin{pmatrix} 2^{28} & 2^{37} \\ 2^{14} & 2^{40} \end{pmatrix} \bmod N\right) \oplus \begin{pmatrix} 27 & 19 \\ 34 & 8 \end{pmatrix} \\ &= \begin{pmatrix} 58 & 51 \\ 60 & 23 \end{pmatrix} \oplus \begin{pmatrix} 27 & 19 \\ 34 & 8 \end{pmatrix} = \begin{pmatrix} 33 & 32 \\ 30 & 31 \end{pmatrix}. \end{aligned}$$

Upon receiving the above ciphertext, Alice decrypts it as follows:

$$\begin{aligned}
 M' &= H(f(p)^3 \cdot c \cdot f(p)^5) \oplus d \\
 &= H\left(\left(\begin{pmatrix} 35 & 12 \\ 63 & 9 \end{pmatrix}^3 \begin{pmatrix} 29 & 40 \\ 52 & 6 \end{pmatrix} \begin{pmatrix} 35 & 12 \\ 63 & 9 \end{pmatrix}^5\right) \oplus \begin{pmatrix} 33 & 32 \\ 30 & 31 \end{pmatrix}\right) \\
 &= H\left(\left(\begin{pmatrix} 28 & 37 \\ 14 & 40 \end{pmatrix}\right) \oplus \begin{pmatrix} 33 & 32 \\ 30 & 31 \end{pmatrix}\right) \\
 &= \left(\left(\begin{pmatrix} 2^{28} & 2^{37} \\ 2^{14} & 2^{40} \end{pmatrix} \bmod N\right) \oplus \begin{pmatrix} 33 & 32 \\ 30 & 31 \end{pmatrix}\right) \\
 &= \begin{pmatrix} 58 & 51 \\ 60 & 23 \end{pmatrix} \oplus \begin{pmatrix} 33 & 32 \\ 30 & 31 \end{pmatrix} = \begin{pmatrix} 27 & 19 \\ 34 & 8 \end{pmatrix} = M.
 \end{aligned}$$

**Example 3.** Encryption/Decryption over matrix rings for the new scheme. Suppose that in the new scheme, the message space is  $\mathcal{M} = \left\{ \begin{pmatrix} a & b \\ c & 0 \end{pmatrix} : a, b, c \in \mathbb{Z}_N \right\}$ , and

$$H_1 : \mathcal{M} \times \mathbb{Z}_N \rightarrow \mathbb{Z}_N^+[x], \left( \begin{pmatrix} a & b \\ c & 0 \end{pmatrix}, r \right) \mapsto 2^r + 2^a x + 2^b x^2 + 2^c x^3,$$

For more simplicity, we define

$$(M \parallel r) \triangleq \begin{pmatrix} a & b \\ c & 0 \end{pmatrix} + \begin{pmatrix} 0 & 0 \\ 0 & r \end{pmatrix} = \begin{pmatrix} a & b \\ c & r \end{pmatrix}.$$

In addition,  $H_2 : M_2(\mathbb{Z}_N) \rightarrow \mathcal{M} \times \mathbb{Z}_N$  can be defined as

$$\begin{pmatrix} a & b \\ c & d \end{pmatrix} \mapsto (M \parallel r),$$

where  $M = \begin{pmatrix} 2^a & 2^b \\ 2^c & 0 \end{pmatrix} \bmod N, r = 2^d \bmod N$ . Suppose that Alice's private/public keys are as the same as that in the basic scheme. Bob picks a message  $M = \begin{pmatrix} 27 & 19 \\ 34 & 0 \end{pmatrix}$ . Suppose  $r = 35$ . Then, he extracts a polynomial  $h(x)$  and compute  $h(p)$  as follows:

$$\begin{aligned}
 h(x) &= 2^{35} + 2^{27}x + 2^{19}x^2 + 2^{34}x^3 \bmod N \\
 &= 32 + 29x + 72x^2 + 16x^3,
 \end{aligned}$$

$$h(p) = 32 \cdot I + 29 \cdot \begin{pmatrix} 2 & 5 \\ 7 & 4 \end{pmatrix} + 72 \cdot \begin{pmatrix} 2 & 5 \\ 7 & 4 \end{pmatrix}^2 + 16 \cdot \begin{pmatrix} 2 & 5 \\ 7 & 4 \end{pmatrix}^3 = \begin{pmatrix} 37 & 30 \\ 42 & 49 \end{pmatrix} \neq \mathbf{0}.$$

Note that if  $h(x)$  does not satisfy the condition  $h(p) \neq \mathbf{0}$ , he should transform  $h(x)$  to

$\tilde{h}(x) = h(x) + \Delta$ , where

$$\Delta = \min \left\{ \delta \in \mathbb{Z}_{\geq 0} : h(p) + \delta \cdot \begin{pmatrix} 1 & 0 \\ 0 & 1 \end{pmatrix} \neq \mathbf{0} \right\}.$$

Then, the ciphertext is  $(c, d)$ , where

$$\begin{aligned} c &= h(p)^3 q h(p)^5 = \begin{pmatrix} 37 & 30 \\ 42 & 49 \end{pmatrix}^3 \begin{pmatrix} 1 & 9 \\ 3 & 2 \end{pmatrix} \begin{pmatrix} 37 & 30 \\ 42 & 49 \end{pmatrix}^5 = \begin{pmatrix} 65 & 37 \\ 35 & 7 \end{pmatrix}, \\ d &= H(h(p)^3 \cdot pk \cdot h(p)^5) \oplus (M \parallel r) \\ &= H \left( \begin{pmatrix} 37 & 30 \\ 42 & 49 \end{pmatrix}^3 \begin{pmatrix} 49 & 53 \\ 42 & 31 \end{pmatrix} \begin{pmatrix} 37 & 30 \\ 42 & 49 \end{pmatrix}^5 \right) \oplus \begin{pmatrix} 27 & 19 \\ 34 & 35 \end{pmatrix} \\ &= H \left( \begin{pmatrix} 21 & 14 \\ 21 & 7 \end{pmatrix} \right) \oplus \begin{pmatrix} 27 & 19 \\ 34 & 35 \end{pmatrix} \\ &= \left( \begin{pmatrix} 2^{21} & 2^{14} \\ 2^{21} & 2^7 \end{pmatrix} \bmod N \right) \oplus \begin{pmatrix} 27 & 19 \\ 34 & 35 \end{pmatrix} \\ &= \begin{pmatrix} 57 & 60 \\ 57 & 51 \end{pmatrix} \oplus \begin{pmatrix} 27 & 19 \\ 34 & 35 \end{pmatrix} = \begin{pmatrix} 34 & 47 \\ 27 & 16 \end{pmatrix}. \end{aligned}$$

Upon receiving the above ciphertext, Alice decrypts it as follows:

$$\begin{aligned} M' &= H(f(p)^3 \cdot c \cdot f(p)^5) \oplus d \\ &= H \left( \begin{pmatrix} 35 & 12 \\ 63 & 9 \end{pmatrix}^3 \begin{pmatrix} 65 & 37 \\ 35 & 7 \end{pmatrix} \begin{pmatrix} 35 & 12 \\ 63 & 9 \end{pmatrix}^5 \right) \oplus \begin{pmatrix} 34 & 47 \\ 27 & 16 \end{pmatrix} \end{aligned}$$

$$\begin{aligned}
&= H \left( \begin{pmatrix} 21 & 14 \\ 21 & 7 \end{pmatrix} \right) \oplus \begin{pmatrix} 34 & 47 \\ 27 & 16 \end{pmatrix} \\
&= \left( \begin{pmatrix} 2^{21} & 2^{14} \\ 2^{21} & 2^7 \end{pmatrix} \bmod N \right) \oplus \begin{pmatrix} 34 & 47 \\ 27 & 16 \end{pmatrix} \\
&= \begin{pmatrix} 57 & 60 \\ 57 & 51 \end{pmatrix} \oplus \begin{pmatrix} 34 & 47 \\ 27 & 16 \end{pmatrix} = \begin{pmatrix} 27 & 19 \\ 34 & 35 \end{pmatrix} \\
&= \begin{pmatrix} 27 & 19 \\ 34 & 0 \end{pmatrix} + \begin{pmatrix} 0 & 0 \\ 0 & 35 \end{pmatrix} = M \parallel r.
\end{aligned}$$

### 5.3.6 $\mathbb{Z}$ -Modular Method over Noncommutative Groups/ Semigroups

Suppose that  $(G, \cdot, 1_G)$  is a noncommutative group,  $(R, +, \cdot, 1_R)$  is a ring and there is monomorphism  $\tau : (G, \cdot, 1_G) \rightarrow (R, \cdot, 1_R)$ . Then, the inverse map  $\tau^{-1} : \tau(G) \rightarrow G$  is also a well-defined monomorphism. For  $a, b \in G$ , if  $\tau(a) + \tau(b) \in \tau(G)$ , we define

$$a \boxplus b \triangleq \tau^{-1}(\tau(a) + \tau(b)), \quad (5.16)$$

and call it the *quasi-sum* of  $a$  and  $b$ . Similarly, for  $k \in R$  and  $a \in G$ , if  $k \cdot \tau(a) \in \tau(G)$ , we define

$$k \boxtimes a \triangleq \tau^{-1}(k \cdot \tau(a)), \quad (5.17)$$

and call it the *k quasi-multiple* of  $a$ .

Given  $a, b \in G$  and  $k \cdot \tau(a) + \tau(b) \in \tau(G)$ , it is easy to see that the monomorphism  $\tau$  is *quasi-linear*, which satisfies that the following equality

$$\begin{aligned}
\tau(k \boxtimes a \boxplus b) &= \tau((k \boxtimes a) \boxplus b) \stackrel{d \leftarrow k \boxtimes a}{=} \tau(d \boxplus b) \\
&= \tau(\tau^{-1}(\tau(d) + \tau(b))) \\
&= \tau(\tau^{-1}(\tau(\tau^{-1}(k \cdot \tau(a))) + \tau(b))) \\
&= \tau(\tau^{-1}(k \cdot \tau(a) + \tau(b))) = k \cdot \tau(a) + \tau(b).
\end{aligned}$$

Furthermore, for  $f(x) = z_0 + z_1x + \cdots + z_nx^n \in \mathbb{Z}[x]$  and  $a \in G$ , if  $f(\tau(a)) = z_0 \cdot 1_R + z_1 \cdot \tau(a) + \cdots + z_n \cdot \tau(a)^n \in \tau(G)$ , we define

$$\hat{f}(a) \triangleq \tau^{-1}(f(\tau(a))) = \tau^{-1}(z_0 \cdot 1_R + z_1 \cdot \tau(a) + \cdots + z_n \cdot \tau(a)^n), \quad (5.18)$$

and call it the *quasi-polynomial* of  $f$  with respect to  $a$ . (Notice that, for arbitrary  $a, b \in G, k \in R$  and  $f(x) \in \mathbb{Z}[x]$ ,  $a \boxplus b, k \boxtimes a$  and  $\hat{f}(a)$  are not always well-defined)

**Theorem 5.3.11.** *Suppose that  $(G, \cdot, 1_G)$  is a noncommutative group,  $(R, +, \cdot, 1_R)$  is a ring and  $\tau : (G, \cdot, 1_G) \rightarrow (R, \cdot, 1_R)$  is a monomorphism. For some  $a \in G$  and some  $f(x), h(x) \in \mathbb{Z}[x]$ , if both  $\hat{f}(a)$  and  $\hat{h}(a)$  are well-defined, we have*

$$\hat{f}(a) \cdot \hat{h}(a) = \hat{h}(a) \cdot \hat{f}(a).$$

*Proof.*

$$\begin{aligned} \hat{f}(a) \cdot \hat{h}(a) &= \tau^{-1}(f(\tau(a))) \cdot \tau^{-1}(h(\tau(a))) \\ &= \tau^{-1}(f(\tau(a)) \cdot h(\tau(a))) \\ &= \tau^{-1}(h(\tau(a)) \cdot f(\tau(a))) \\ &= \tau^{-1}(h(\tau(a))) \cdot \tau^{-1}(f(\tau(a))) \\ &= \hat{h}(a) \cdot \hat{f}(a). \end{aligned}$$

Now, similar to the discussion in the subsection 5.3.2, we can discuss new problems over the noncommutative group  $(G, \cdot, 1_G)$ . Suppose that  $(R, +, \cdot, 1_R)$  is a ring and  $\tau : (G, \cdot, 1_G) \rightarrow (R, \cdot, 1_R)$  is a monomorphism. For any randomly picked element  $a \in G$ , we define a set  $P_a \subseteq G$  by

$$P_a \triangleq \{\hat{f}(a) : f(x) \in \mathbb{Z}[x], f(\tau(a)) \in \tau(G)\}.$$

Then, we can define the PSD and PDH problems over  $(G, \cdot)$  by a similar way:

- **Polynomial Symmetrical Decomposition (PSD\*) Problem:** Given  $(a, x, y) \in G^3$  and  $m, n \in \mathbb{Z}$ , find  $z \in P_a$  so that  $y = z^m x z^n$ , where  $G$  is a noncommutative group.
- **Polynomial Diffie–Hellman (PDH\*) Problem:** Given  $a, x, x^{z_1}, x^{z_2}, a, x \in G$  and  $z_1, z_2 \in P_a$ , compute  $x^{z_1 z_2}$  or  $x^{z_2 z_1}$ , where  $G$  is a noncommutative group.

Accordingly, the PSD\* (PDH\*, respectively) cryptographic assumptions over  $(G, \cdot)$  says that PSD\* (PDH\*, respectively) problems over  $(G, \cdot)$  is intractable, i.e., there does not exist probabilistic polynomial time algorithm which can solve PSD\* (PDH\*, respectively) problems over  $(G, \cdot)$  with non-negligible advantage with respect to the parameters of the problem scale.

Under these assumptions, we have

**Protocol 5.3.12.** Diffie–Hellman-Like Key Agreement Protocol over Noncommutative Group:

- (0) Suppose that  $(G, \cdot, 1_G)$  is a noncommutative group,  $(R, +, \cdot, 1_R)$  is a ring and  $\tau : (G, \cdot, 1_G) \rightarrow (R, +, \cdot, 1_R)$  is a monomorphism. Alice sends two random small, positive integers (say, less than 10)  $m, n \in \mathbb{Z}$  and two random elements  $a, b \in G$  to Bob.
- (1) Alice chooses  $f(x) \in \mathbb{Z}[x]$  at random so that  $\hat{f}(a)$  is well-defined, i.e.,  $f(\tau(a)) \in \tau(G)$ . Then, Alice takes  $\hat{f}(a)$  as her private key.
- (2) Bob chooses  $h(x) \in \mathbb{Z}[x]$  at random so that  $\hat{h}(a)$  is well-defined, i.e.,  $h(\tau(a)) \in \tau(G)$ . Then, Bob takes  $\hat{h}(a)$  as his private key.
- (3) Alice computes  $r_A = \hat{f}(a)^m \cdot b \cdot \hat{f}(a)^n$  and sends  $r_A$  to Bob.
- (4) Bob computes  $r_B = \hat{h}(a)^m \cdot b \cdot \hat{h}(a)^n$  and sends  $r_B$  to Alice.



(5) Alice computes  $K_A = \hat{f}(a)^m \cdot r_B \cdot \hat{f}(a)^n$  as the shared session key.

(6) Bob computes  $K_B = \hat{h}(a)^m \cdot r_A \cdot \hat{h}(a)^n$  as the shared session key.

In practice, the steps (0), (1), and (3) can be finished simultaneously and require only one pass communication from Alice to Bob. After that, the steps (2) and (4) can be finished simultaneously and require another pass communication from Bob to Alice. Finally, Alice and Bob can execute the steps (5) and (6), respectively. We can depict the protocol in Figure 5.5.

Pass	Alice	Bob
	$m, n \xleftarrow{\$} \mathbb{Z}$ $a, b \xleftarrow{\$} G$ $f(x) \xleftarrow{\$} \mathbb{Z}[x]$ s.t. $f(\tau(a)) \in \tau(G)$	
1	$\xrightarrow{m, n, a, b, \hat{f}(a)^m b \hat{f}(a)^n}$	
	$h(g) \xleftarrow{\$} \mathbb{Z}[x]$ s.t. $h(\tau(a)) \in \tau(G)$	
2	$\xleftarrow{\hat{h}(a)^m b \hat{h}(a)^n}$	
	$K_A = \hat{f}(a)^m \hat{h}(a)^m b \hat{h}(a)^n \hat{f}(a)^n = K_B = \hat{h}(a)^m \hat{f}(a)^m b \hat{f}(a)^n \hat{h}(a)^n$	

Figure 5.5: Diffie–Hellman-Like Key Agreement over Noncommutative Groups

Similarly, it is easy to describe ElGamal-like encryption schemes, including the basic scheme and the improved scheme as well, by using noncommutative groups as the underlying algebraic structure.

**Protocol 5.3.13.** ElGamal-Like Encryption (Basic) Scheme over Noncommutative Group:

- **Setup:** Suppose that  $(G, \cdot, 1_G)$  is a noncommutative group,  $(R, +, \cdot, 1_R)$  is a ring and  $\tau : (G, \cdot, 1_G) \rightarrow (R, \cdot, 1_R)$  is a monomorphism. We assume that SDP on  $G$  is intractable. Pick two small positive integers  $m, n \in \mathbb{Z}$  and two elements  $p, q \in G$  at random. Let  $H : G \rightarrow \mathcal{M}$  be a cryptographic hash function which maps  $G$

to the message space  $\mathcal{M}$ . Then, set the tuple  $\langle G, R, \tau, m, n, p, q, \mathcal{M}, H \rangle$  as the public parameters of the system.

- **Key generation:** Each user chooses a random polynomial  $f(x) \in \mathbb{Z}[x]$  so that  $f(\tau(p)) \in \tau(G)$  and takes  $sk \triangleq \hat{f}(p)$  as his private key, then computes and publishes his public key  $pk \triangleq \hat{f}(p)^m \cdot q \cdot \hat{f}(p)^n \in G$ .
- **Encryption:** Given a message  $M \in \mathcal{M}$  and receiver's key  $pk \in G$ , the sender chooses a random polynomial  $h(x) \in \mathbb{Z}[x]$  so that  $h(\tau(p)) \in \tau(G)$ . Compute

$$c = \hat{h}(p)^m \cdot q \cdot \hat{h}(p)^n, \quad d = H(\hat{h}(p)^m \cdot pk \cdot \hat{h}(p)^n) \oplus M,$$

and output the ciphertext  $(c, d) \in G \times \mathcal{M}$ .

- **Decryption:** Upon receiving a ciphertext  $(c, d) \in R \times \mathcal{M}$ , the receiver, by using his private key  $\hat{f}(p)$ , computes the plaintext

$$M = H(\hat{f}(p)^m \cdot c \cdot \hat{f}(p)^n) \oplus d$$

**Protocol 5.3.14.** Improved ElGamal-Like Encryption Scheme over Noncommutative Group:

- **Setup:** The parameters  $G, R, \tau, m, n, p, q, \mathcal{M}$  are set as that in Protocol 5.3.13. In addition, suppose that  $H_1 : \{0, 1\}^{k+k_0} \rightarrow \mathbb{Z}^+[x]$  and  $H_2 : G \rightarrow \{0, 1\}^{k+k_0}$  are two cryptographic hash functions, where  $k$  is the length of a message and  $k_0$  is the length of random string for padding.
- **Key generation:** See Protocol 5.3.13.
- **Encryption:** Given a message  $M \in \mathcal{M}$  and receiver's key  $pk \in G$ , the sender chooses a random  $r \in \{0, 1\}^{k_0}$  and *extracts* a polynomial  $h(x) \in \mathbb{Z}[x]$  so that

$h(\tau(p)) \in \tau(G)$  and then computes

$$c = \hat{h}(p)^m \cdot q \cdot \hat{h}(p)^n, \quad d = H_2(\hat{h}(p)^m \cdot pk \cdot \hat{h}(p)^n) \oplus (M \parallel r),$$

and outputs the ciphertext  $(c, d) \in R \times \{0, 1\}^{k+k_0}$ .

- **Decryption:** Upon receiving a ciphertext  $(c, d) \in R \times \{0, 1\}^{k+k_0}$ , the receiver, by using his private key  $\hat{f}(p)$ , computes

$$M' = H_2(\hat{f}(p)^m \cdot c \cdot \hat{f}(p)^n) \oplus d.$$

Extract  $g(x) = H_1(M') \in \mathbb{Z}[x]$  and check whether  $c = \hat{g}(p)^m \cdot q \cdot \hat{g}(p)^n$  holds.

If so, output the first  $k$  bits of  $M'$ ; otherwise, output  $\perp$ .

The security and related proofs of the above cryptosystems are very similar, except replacing the PDH assumption over noncommutative ring  $R$  with the PDH\* assumption over noncommutative group  $G$ . The similar constructions can also be obtained for the case when  $G$  is a noncommutative semigroup, except replacing  $\mathbb{Z}$  and  $\mathbb{Z}[x]$  with  $\mathbb{Z}^+$  and  $\mathbb{Z}^+[x]$  respectively. The main differences between the protocols and schemes over noncommutative rings and semigroups and those over noncommutative groups are:

- (1)  $m, n \in \mathbb{Z}^+$  for the schemes over rings and semigroups, while  $m, n \in \mathbb{Z}$  for the schemes over groups;
- (2)  $f, h \in \mathbb{Z}^+[x]$  for the schemes over rings and semigroups, while  $f, h \in \{g \in \mathbb{Z}[x] : g(\tau(a)) \in \tau(G)\} \subseteq \mathbb{Z}[x]$  for the schemes over groups.

### 5.3.7 Instantiation and Illustration (II)

To illustrate our method, we take the minimal noncommutative group, say, the symmetric group  $S_3$ , as the underlying algebraic structure. We choose  $M_2(\mathbb{Z}_2)$  as the re-

quired ring. Define the required monomorphism  $\tau : S_3 \rightarrow M_2(\mathbb{Z}_2)$  as follows:

$$\begin{aligned} \begin{pmatrix} 1 & 2 & 3 \\ 1 & 2 & 3 \end{pmatrix} &\mapsto \begin{pmatrix} 1 & 0 \\ 0 & 1 \end{pmatrix}, & \begin{pmatrix} 1 & 2 & 3 \\ 1 & 3 & 2 \end{pmatrix} &\mapsto \begin{pmatrix} 1 & 1 \\ 0 & 1 \end{pmatrix}, & \begin{pmatrix} 1 & 2 & 3 \\ 2 & 1 & 3 \end{pmatrix} &\mapsto \begin{pmatrix} 0 & 1 \\ 1 & 0 \end{pmatrix}, \\ \begin{pmatrix} 1 & 2 & 3 \\ 2 & 3 & 1 \end{pmatrix} &\mapsto \begin{pmatrix} 0 & 1 \\ 1 & 1 \end{pmatrix}, & \begin{pmatrix} 1 & 2 & 3 \\ 3 & 1 & 2 \end{pmatrix} &\mapsto \begin{pmatrix} 1 & 1 \\ 1 & 0 \end{pmatrix}, & \begin{pmatrix} 1 & 2 & 3 \\ 3 & 2 & 1 \end{pmatrix} &\mapsto \begin{pmatrix} 1 & 0 \\ 1 & 1 \end{pmatrix}. \end{aligned}$$

**Example 4. Diffie–Hellman-Like Key Agreement over  $S_3$**

Alice chooses

$$m = 3, n = 5, A = \begin{pmatrix} 1 & 2 & 3 \\ 2 & 3 & 1 \end{pmatrix}, B = \begin{pmatrix} 1 & 2 & 3 \\ 2 & 1 & 3 \end{pmatrix},$$

and picks a polynomial  $f(x) = 4x^2 + x + 2$ . Clearly,  $f(A) \in \tau(S_3)$ . Then, Alice computes

$$\begin{aligned} f(A) &= \tau^{-1}(f(\tau(A))) \\ &= \tau^{-1}\left(4 \cdot \begin{pmatrix} 0 & 1 \\ 1 & 1 \end{pmatrix}^2 + \begin{pmatrix} 0 & 1 \\ 1 & 1 \end{pmatrix} + 2 * I\right) \\ &= \tau^{-1}\left(\begin{pmatrix} 0 & 1 \\ 1 & 1 \end{pmatrix}\right) = \begin{pmatrix} 1 & 2 & 3 \\ 2 & 3 & 1 \end{pmatrix} \\ r_A &= \begin{pmatrix} 1 & 2 & 3 \\ 2 & 3 & 1 \end{pmatrix}^3 \circ \begin{pmatrix} 1 & 2 & 3 \\ 2 & 1 & 3 \end{pmatrix} \circ \begin{pmatrix} 1 & 2 & 3 \\ 2 & 3 & 1 \end{pmatrix}^5 = \begin{pmatrix} 1 & 2 & 3 \\ 3 & 2 & 1 \end{pmatrix}. \end{aligned}$$

Send  $m, n, A, B$ , and  $r_A$  to Bob.

Upon receiving  $m, n, A, B$ , and  $f(A)$  from Alice, Bob chooses another random polynomial  $h(x) = 4x^4 + x^3 + 4x^2 + 3x + 4$ . Clearly,  $h(A) \in \tau(S_3)$ . Then, Bob

computes

$$\begin{aligned}
 h(A) &= \tau^{-1}(h(\tau(A))) \\
 &= \tau^{-1}\left(4 \cdot \begin{pmatrix} 0 & 1 \\ 1 & 1 \end{pmatrix}^4 + \begin{pmatrix} 0 & 1 \\ 1 & 1 \end{pmatrix}^3 + 4 \cdot \begin{pmatrix} 0 & 1 \\ 1 & 1 \end{pmatrix}^2 + 3 \cdot \begin{pmatrix} 0 & 1 \\ 1 & 1 \end{pmatrix} + 4 * I\right) \\
 &= \tau^{-1}\left(\begin{pmatrix} 1 & 1 \\ 1 & 0 \end{pmatrix}\right) = \begin{pmatrix} 1 & 2 & 3 \\ 3 & 1 & 2 \end{pmatrix}, \\
 r_B &= \begin{pmatrix} 1 & 2 & 3 \\ 3 & 1 & 2 \end{pmatrix}^3 \circ \begin{pmatrix} 1 & 2 & 3 \\ 2 & 1 & 3 \end{pmatrix} \circ \begin{pmatrix} 1 & 2 & 3 \\ 3 & 1 & 2 \end{pmatrix}^5 = \begin{pmatrix} 1 & 2 & 3 \\ 1 & 3 & 2 \end{pmatrix}.
 \end{aligned}$$

Send  $r_B$  to Alice.

Finally, Alice extracts the session key

$$K_A = \begin{pmatrix} 1 & 2 & 3 \\ 2 & 3 & 1 \end{pmatrix}^3 \circ \begin{pmatrix} 1 & 2 & 3 \\ 1 & 3 & 2 \end{pmatrix} \circ \begin{pmatrix} 1 & 2 & 3 \\ 2 & 3 & 1 \end{pmatrix}^5 = \begin{pmatrix} 1 & 2 & 3 \\ 2 & 1 & 3 \end{pmatrix},$$

and Bob extracts the session key

$$K_B = \begin{pmatrix} 1 & 2 & 3 \\ 3 & 1 & 2 \end{pmatrix}^3 \circ \begin{pmatrix} 1 & 2 & 3 \\ 3 & 2 & 1 \end{pmatrix} \circ \begin{pmatrix} 1 & 2 & 3 \\ 3 & 1 & 2 \end{pmatrix}^5 = \begin{pmatrix} 1 & 2 & 3 \\ 2 & 1 & 3 \end{pmatrix}.$$

### Example 5. Encryption/Decryption over $S_3$

We can implement encryption/decryption by ways which are very similar to those in Example 2 and Example 3.

At first, let us choose two prime  $p$  and  $q$  so that  $q|p-1$  and set  $Z_p$  as the message space  $\mathcal{M}$ . Also, we assume that  $g$  is a generator of order  $q$ . Then, we define

$$H : S_3 \rightarrow \mathcal{M}, \begin{pmatrix} 1 & 2 & 3 \\ \sigma_1 & \sigma_2 & \sigma_3 \end{pmatrix} \mapsto g^{\sigma_1 + 2 \cdot \sigma_2 + 2^2 \cdot \sigma_3} \bmod p,$$

$$H_1 : \mathcal{M} \times \mathbb{Z}_N \rightarrow \mathbb{Z}^+[x], (M, r) \mapsto r_0 + r_1x + r_2x^2 + \cdots + r_nx^n,$$

where all coefficients  $r_i, 0 \leq i \leq k$ , are determined by the following process<sup>8</sup>:

$$\begin{aligned}
 M &= rk_0 + r_0, \quad 0 < r_0 < r, \\
 r &= r_0k_1 + r_1, \quad 0 < r_1 < r_0, \\
 r_0 &= r_1k_2 + r_2, \quad 0 < r_2 < r_1, \\
 \dots &\quad \dots \dots \dots \dots \dots \dots \\
 r_{l-2} &= r_{l-1}k_l + r_l, \quad 0 < r_l < r_{l-1}, \\
 \dots &\quad \dots \dots \dots \dots \dots \dots \\
 r_{n-1} &= r_nk_{n+1} + r_{n+1}, \quad r_{n+1} = 0.
 \end{aligned}$$

For simplicity, we define

$$M \parallel r \triangleq r \cdot p + M \in \mathbb{Z}.$$

It is easy to extract  $M$  from  $M \parallel r$  by  $M = (M \parallel r) \bmod p$ . Another required hash  $H_2 : S_3 \rightarrow \mathcal{M} \times \mathbb{Z}_N$  can be defined as

$$\begin{pmatrix} 1 & 2 & 3 \\ \sigma_1 & \sigma_2 & \sigma_3 \end{pmatrix} \mapsto (g^{\sigma_1+2 \cdot \sigma_2+2^2 \cdot \sigma_3} \bmod p, g^{\sigma_3+2 \cdot \sigma_1+2^2 \cdot \sigma_2} \bmod p).$$

Next, let  $p = 23, q = 11$  and  $g = 2$ . Suppose that other system parameters are

$$m = 3, n = 5, p = \begin{pmatrix} 1 & 2 & 3 \\ 2 & 3 & 1 \end{pmatrix}, q = \begin{pmatrix} 1 & 2 & 3 \\ 2 & 1 & 3 \end{pmatrix}.$$

Suppose that the polynomial  $f(x)$  picked by Alice is just that in example 4. Then, Alice's private key is  $\hat{f}(p) = \begin{pmatrix} 1 & 2 & 3 \\ 2 & 3 & 1 \end{pmatrix}$ , and the corresponding public key is

$$pk \triangleq \hat{f}(p)^3 q \hat{f}(p)^5 = \begin{pmatrix} 1 & 2 & 3 \\ 2 & 3 & 1 \end{pmatrix}^3 \circ \begin{pmatrix} 1 & 2 & 3 \\ 2 & 1 & 3 \end{pmatrix} \circ \begin{pmatrix} 1 & 2 & 3 \\ 2 & 3 & 1 \end{pmatrix}^5 = \begin{pmatrix} 1 & 2 & 3 \\ 3 & 2 & 1 \end{pmatrix}.$$

---

<sup>8</sup>Here, we assume that  $\gcd(M, r) \neq r$ ; Otherwise, we set  $r = g^r \bmod p$  and then resume the process. Also, we assume that  $M > r$ ; Otherwise, we can swap them in advance.

Bob picks a message  $M = 17$  and a polynomial  $h(x) = 4x^4 + x^3 + 4x^2 + 3x + 4$ .

Then, the permutation  $\hat{h}(p) = \begin{pmatrix} 1 & 2 & 3 \\ 3 & 1 & 2 \end{pmatrix}$ . Compute the ciphertext  $(c, d)$  as follows:

$$\begin{aligned}
 c &= \hat{h}(p)^3 q \hat{h}(p)^5 = \begin{pmatrix} 1 & 2 & 3 \\ 3 & 1 & 2 \end{pmatrix}^3 \circ \begin{pmatrix} 1 & 2 & 3 \\ 2 & 1 & 3 \end{pmatrix} \circ \begin{pmatrix} 1 & 2 & 3 \\ 3 & 1 & 2 \end{pmatrix}^5 = \begin{pmatrix} 1 & 2 & 3 \\ 1 & 3 & 2 \end{pmatrix}, \\
 d &= H(\hat{h}(p)^3 \cdot pk \cdot \hat{h}(p)^5) \oplus M \\
 &= H\left(\begin{pmatrix} 1 & 2 & 3 \\ 3 & 1 & 2 \end{pmatrix}^3 \circ \begin{pmatrix} 1 & 2 & 3 \\ 3 & 2 & 1 \end{pmatrix} \circ \begin{pmatrix} 1 & 2 & 3 \\ 3 & 1 & 2 \end{pmatrix}^5\right) \oplus 17 \\
 &= H\left(\begin{pmatrix} 1 & 2 & 3 \\ 2 & 1 & 3 \end{pmatrix}\right) \oplus 17 \\
 &= \left(2^2 + 2^{2 \cdot 1} + 2^{2^2 \cdot 3} \bmod 23\right) \oplus 17 \\
 &= 18 \oplus 17 = 3.
 \end{aligned}$$

On receiving the ciphertext, Alice decrypts it as follows:

$$\begin{aligned}
 M' &= H(\hat{f}(p)^3 \cdot c \cdot \hat{f}(p)^5) \oplus d \\
 &= H\left(\begin{pmatrix} 1 & 2 & 3 \\ 2 & 3 & 1 \end{pmatrix}^3 \circ \begin{pmatrix} 1 & 2 & 3 \\ 1 & 3 & 2 \end{pmatrix} \circ \begin{pmatrix} 1 & 2 & 3 \\ 2 & 3 & 1 \end{pmatrix}^5\right) \oplus 3 \\
 &= H\left(\begin{pmatrix} 1 & 2 & 3 \\ 2 & 1 & 3 \end{pmatrix}\right) \oplus 3 = \left(2^2 + 2^{2 \cdot 1} + 2^{2^2 \cdot 3} \bmod 23\right) \oplus 3 = 18 \oplus 3 = 17 = M.
 \end{aligned}$$

For the improved encryption, Bob picks  $M = 19$  and sets  $r = 7$ . Then, he extracts a polynomial as follows:

$$h(x) \triangleq H_1(19, 7) = 5 + 2x + x^2 + x^3.$$

Clearly,  $h(\tau(p)) \in \tau(S_3)$ . Note that if  $h(x)$  does not satisfy the condition of  $h(\tau(p)) \in \tau(S_3)$ , he should transform  $h(x)$  to  $\tilde{h}(x) = h(x) + \Delta$ , where

$$\Delta = \min \left\{ \delta \in \mathbb{Z}_{\geq 0} : h(\tau(p)) + \delta \cdot \begin{pmatrix} 1 & 0 \\ 0 & 1 \end{pmatrix} \in \tau(S_3) \right\}.$$

Then,

$$\begin{aligned}
 \hat{h}(p) &= \tau^{-1}(h(\tau(p))) \\
 &= \tau^{-1}\left(5 \cdot \begin{pmatrix} 1 & 0 \\ 0 & 1 \end{pmatrix} + 2 \cdot \begin{pmatrix} 0 & 1 \\ 1 & 1 \end{pmatrix} + \begin{pmatrix} 0 & 1 \\ 1 & 1 \end{pmatrix}^2 + \begin{pmatrix} 0 & 1 \\ 1 & 1 \end{pmatrix}^3\right) \\
 &= \tau^{-1}\left(\begin{pmatrix} 1 & 1 \\ 0 & 1 \end{pmatrix}\right) = \begin{pmatrix} 1 & 2 & 3 \\ 3 & 1 & 2 \end{pmatrix}.
 \end{aligned}$$

Consequently, the corresponding ciphertext  $(c, d)$  is

$$\begin{aligned}
 c &= \hat{h}(p)^3 q \hat{h}(p)^5 = \begin{pmatrix} 1 & 2 & 3 \\ 3 & 1 & 2 \end{pmatrix}^3 \circ \begin{pmatrix} 1 & 2 & 3 \\ 2 & 1 & 3 \end{pmatrix} \circ \begin{pmatrix} 1 & 2 & 3 \\ 3 & 1 & 2 \end{pmatrix}^5 = \begin{pmatrix} 1 & 2 & 3 \\ 1 & 3 & 2 \end{pmatrix}, \\
 d &= H(\hat{h}(p)^3 \cdot pk \cdot \hat{h}(p)^5) \oplus (M \parallel r) \\
 &= H\left(\begin{pmatrix} 1 & 2 & 3 \\ 3 & 1 & 2 \end{pmatrix}^3 \circ \begin{pmatrix} 1 & 2 & 3 \\ 3 & 2 & 1 \end{pmatrix} \circ \begin{pmatrix} 1 & 2 & 3 \\ 3 & 1 & 2 \end{pmatrix}^5\right) \oplus (19 + 5 \cdot 23) \\
 &= H\left(\begin{pmatrix} 1 & 2 & 3 \\ 2 & 1 & 3 \end{pmatrix}\right) \oplus 134 \\
 &= \left(2^2 + 2^{2 \cdot 1} + 2^{2^2 \cdot 3} \bmod 23\right) \oplus 134 = 18 \oplus 134 = 148.
 \end{aligned}$$

On receiving the ciphertext, Alice decrypts it as follows:

$$\begin{aligned}
 M' &= H(f(p)^3 \cdot c \cdot f(p)^5) \oplus d \\
 &= H\left(\begin{pmatrix} 1 & 2 & 3 \\ 2 & 3 & 1 \end{pmatrix}^3 \circ \begin{pmatrix} 1 & 2 & 3 \\ 1 & 3 & 2 \end{pmatrix} \circ \begin{pmatrix} 1 & 2 & 3 \\ 2 & 3 & 1 \end{pmatrix}^5\right) \oplus 148 \\
 &= H\left(\begin{pmatrix} 1 & 2 & 3 \\ 2 & 1 & 3 \end{pmatrix}\right) \oplus 3 \\
 &= \left(2^2 + 2^{2 \cdot 1} + 2^{2^2 \cdot 3} \bmod 23\right) \oplus 148 = 18 \oplus 148 = 134.
 \end{aligned}$$

Then,  $M = M' \bmod 23 = 19$ .



## 5.4 Using Monomials in $\mathbb{Z}$ -Modular Method

If the polynomials used in  $\mathbb{Z}$ -modular method are constrained to be monomials, then the conjugacy related assumptions can be viewed as special cases of the assumptions defined in  $\mathbb{Z}$ -modular method. Under these assumptions, we [287] recently proposed several noncommutative cryptosystems.

### 5.4.1 Conjugate Left Self-Distributed System (Conj-LD)

At first, let us recall the definition of the conjugacy search problem and the so-called left self-distributive system [94].

Let  $(G, \circ, 1)$  be a noncommutative monoid with identity 1. For  $a \in G$ , if there exists an element  $b \in G$  so that  $a \circ b = 1 = b \circ a$ , then we say that  $a$  is *invertible*, and call  $b$  an *inverse of  $a$* . Note that not all elements in  $G$  are invertible. If the inverse of  $a$ , exists, it is unique, and denoted by  $a^{-1}$ . In a monoid, one can define positive integer powers of an element  $a$ :  $a^1 = a$ , and  $a^n = \underbrace{a \circ \dots \circ a}_{n \text{ times}}$  for  $n > 1$ . If  $b$  is the inverse of  $a$ , one can also define negative powers of  $a$  by setting  $a^{-1} = b$  and  $a^{-n} = \underbrace{b \circ \dots \circ b}_{n \text{ times}}$  for  $n > 1$ . In addition, let us denote  $a^0 = 1$  for each element  $a \in G$ . Throughout this section, let  $G^{-1}$  be the set of all invertible elements in  $G$ , i.e.,

$$G^{-1} \triangleq \{a \in G : \exists b \in G \text{ so that } a \circ b = 1 = b \circ a\}. \quad (5.19)$$

In fact,  $(G^{-1}, \circ, 1)$  forms a group. For clarity, we omit “ $\circ$ ” in the following presentation, i.e., writing  $a \circ b$  as  $ab$  directly.

The conjugacy problem is extensively studied in *group theory*. But in this section, we would like to extend the conjugacy concept to monoids by a similar manner: Given a monoid  $G$ , for  $\forall a \in G$  and  $\forall x \in G^{-1}$ ,  $axa^{-1}$  is *conjugate* to  $a$ , and call  $x$  a *conjugator* of the pair  $(a, axa^{-1})$ .

**Definition 5.4.1** (Conjugacy Search Problem, CSP). Let  $G$  be a noncommutative monoid. Given two elements  $a, b \in G$  so that  $b = xax^{-1}$  for some unknown element  $x \in G^{-1}$ , the objective of the conjugacy search problem in  $G$  is to find  $x' \in G^{-1}$  so that  $b = x'ax'^{-1}$  holds. Here,  $x'$  is not required to be  $x$ .

**Definition 5.4.2** (Left self-distributive system, LD [94]). Suppose that  $S$  is a non-empty set,  $F : S \times S \rightarrow S$  is a well-defined function and let us denote  $F(a, b)$  by  $F_a(b)$ . If the following rewritten formula holds,

$$F_r(F_s(p)) = F_{F_r(s)}(F_r(p)), \quad (\forall p, r, s \in S) \quad (5.20)$$

then, we call  $F(\cdot)$  a *left self-distributive system*, abbreviated as LD system.

The terminology “left self-distributive” arises from the following analogical observation: If we consider  $F_r(s)$  as a binary operation  $r * s$ , then the formula (5.20) becomes

$$r * (s * p) = (r * s) * (r * p), \quad (5.21)$$

i.e., the operation “ $*$ ” is left distributive with respect to itself [94].

Combining the above two concepts, one can define the following LD system, named as Conj-LD system.

**Theorem 5.4.3** (Conj-LD System). *Let  $G$  be a noncommutative monoid. The binary function  $F$ , given by the following conjugate operation*

$$F : G^{-1} \times G \rightarrow G, \quad (a, b) \mapsto aba^{-1}, \quad (5.22)$$

*is an LD system, abbreviated as Conj-LD.*

*Proof.* It is easy to see that  $F$  satisfies the rewritten formula (5.20). Thus, the above definitions  $F_a(b)$  is an LD system.  $\square$

In what follows, we prove some properties of the Conj-LD system. These properties are useful from the cryptographic viewpoint, providing that the related hardness assumptions hold.

**Proposition 5.4.4.** *Let  $F$  be a Conj-LD system defined over a noncommutative monoid  $G$ . Given  $a \in G^{-1}$  and  $b, c \in G$ , we have*

$$(i) \ F_a(a) = a;$$

$$(ii) \ F_a(b) = c \Leftrightarrow F_{a^{-1}}(c) = b;$$

$$(iii) \ F_a(bc) = F_a(b)F_a(c).$$

*Proof.* Firstly, since  $aaa^{-1} = a$ , we have  $F_a(a) = a$ . Next,  $F_a(b) = c \Leftrightarrow c = aba^{-1} \Leftrightarrow a^{-1}ca = b \Leftrightarrow F_{a^{-1}}(c) = b$ . Finally,  $F_a(bc) = a(bc)a^{-1} = (aba^{-1})(aca^{-1}) = F_a(b)F_a(c)$ .  $\square$

**Proposition 5.4.5 (Power Law).** *Let  $F$  be a Conj-LD system defined over a noncommutative monoid  $G$ . Suppose that  $a \in G^{-1}$  and  $b \in G$ . Then, for arbitrary three positive integers  $m, s, t$  so that  $m = s + t$ , we have*

$$F_a(b^m) = F_a(b^s)F_a(b^t) = F_a^m(b) \quad \text{and} \quad F_{a^m}(b) = F_{a^s}(F_{a^t}(b)). \quad (5.23)$$

*Proof.* See the property (iii) in Proposition 1 and the definition of the Conj-LD system.

**Remark 5.4.6.** The term of left self-distributive system is due to the following reasons:

- The binary function defined by the formula (5.22) *does* satisfy the left self-distributiveness defined by the formula (5.20). In addition, it is more convenient to use  $F_a(b)$  than to use  $aba^{-1}$  in the sequel presentation, especially when some additional operations are exerted on  $a$  and  $a^{-1}$  simultaneously from both sides.

- In 2006, Dehornoy [94] proposed an authentication scheme based on left self-distributive systems in braid groups. Although some cryptanalysis on Dehornoy's authentication scheme were reported [202], we find that Dehornoy's work is still meaningful at least in the following two aspects: (1) self-distributive systems can be defined over arbitrary noncommutative monoids, rather than braid groups; (2) self-distributive systems have the potential for building variety of cryptographic schemes, rather than authentication schemes. Therefore, we would like to use the terminology of (left) self-distributive system to give Dehornoy the credit for coining this concept.

Now, using the notation of  $F(\cdot)$ , the intractability assumption of the CSP problem in  $G$  can be re-formulated as follows: It is hard to retrieve  $a'$  from the given pair  $(a, F_a(b))$  so that  $F_a(b) = F_{a'}(b)$ . We must take care of the relationship between the intractability assumption of CSP and the hardness CSP instances. The CSP problem is defined here as a *worst-case* problem, whereas from the cryptographic viewpoint, one needs *average-case* hardness for CSP instances. Therefore, we need a practical sampling algorithm that can produce hard CSP instances over  $G$ . For a generic noncommutative monoid  $G$  without explicit definition or presentation, it is difficult to determine whether we can sample hard CSP instances over  $G$ . Afterwards, the CSP instances used in our proposal are always assumed to be hard.

### 5.4.2 New Assumptions in Conj-LD Systems

Let  $G$  be a noncommutative monoid. Given a Conj-LD system  $F$  over a noncommutative monoid  $G$  (cf. Definition 5.4.3). Let  $a \in G^{-1}$  and  $G_a$  denote the subgroup generated by  $\{a, a^{-1}\}$ , i.e.,  $G_a \triangleq \langle a, a^{-1} \rangle$ . Now, given  $b \in G$ , we define the following notations:

- $\mathbb{T} \triangleq \{1, \dots, n\}$  is a *finite* subset of  $\mathbb{Z}$ , where  $n$  is the order of  $G_a$ <sup>9</sup>.

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<sup>9</sup>In case  $G_a$  is *infinite*, we can set  $n$  as a fixed integer that is large enough, say 160 bits, to resist exhaustive

- The symbols “ $\in \mathbb{T}$ ” and “ $\xleftarrow{\$} \mathbb{T}$ ” always indicate sampling procedures that pick random integers uniformly from  $\mathbb{T}$ .
- $\mathcal{K}_{[a,b]} \triangleq \{F_{a^i}(b) : i \in \mathbb{T}\}$  is a *finite* subset of  $G$ .

Note that in cryptographic applications,  $a, b$  should be chosen appropriately so that  $n$  and  $\mathcal{K}_{[a,b]}$  are large enough to resist exhaustive attacks.

**Definition 5.4.7** (CSP-based Computational Diffie–Hellman Problem: CSP-CDH). Let  $F$  be a Conj-LD system over a noncommutative monoid  $G$  and let  $\mathcal{A}$  be an adversary. For given  $a \in G^{-1}$  and  $b \in G$ , consider the following experiment:

Experiment  $\mathbf{Exp}_{F,\mathcal{A}}^{csp-cdh}$

$$\begin{aligned} & i \xleftarrow{\$} \mathbb{T}; X \leftarrow F_{a^i}(b); \\ & j \xleftarrow{\$} \mathbb{T}; Y \leftarrow F_{a^j}(b); \\ & Z \leftarrow \mathcal{A}(X, Y); \\ & \text{if } Z = F_{a^{i+j}}(b) \text{ then } b \leftarrow 1 \text{ else } b \leftarrow 0; \\ & \text{return } b. \end{aligned}$$

Now define the *advantage* of  $\mathcal{A}$  violating the CSP-based computational Diffie–Hellman assumption as

$$\mathbf{Adv}_{F,\mathcal{A}}^{csp-cdh} = \Pr[\mathbf{Exp}_{F,\mathcal{A}}^{csp-cdh} = 1]. \quad (5.24)$$

The CSP-CDH assumption states, roughly, that given  $F_{a^i}(b)$  and  $F_{a^j}(b)$ , where  $i, j$  are drawn at random from  $\mathbb{T}$ , it is hard to compute  $F_{a^{i+j}}(b)$ . Suppose the CSP-CDH assumption is moderate, the adversary is possible to compute something interesting, such as the most significant bit of  $F_{a^{i+j}}(b)$  given  $F_{a^i}(b)$  and  $F_{a^j}(b)$ . This means the CSP-CDH problem is too weak to be directly used in cryptographic applications. Thus, we need the following stronger variants.

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attacks. Furthermore, to ensure the randomness of sampling from  $\mathcal{K}_{[a,b]}$ ,  $n$  should be the order of a factor-group of  $G_a$  modulo the centralizer of  $b$  in  $G_a$ .

**Definition 5.4.8** (CSP-based Decisional Diffie–Hellman Problem: CSP-DDH). Let  $F$  be a Conj-LD system over a noncommutative monoid  $G$  and let  $\mathcal{A}$  be an adversary. Given  $a \in G^{-1}$  and  $b \in G$ , consider the following two experiments in a parallel manner

Experiment $\mathbf{Exp}_{F,\mathcal{A}}^{csp-ddh-real}$	Experiment $\mathbf{Exp}_{F,\mathcal{A}}^{csp-ddh-rand}$
$i \xleftarrow{\$} \mathbb{T}; X \leftarrow F_{a^i}(b);$	$i \xleftarrow{\$} \mathbb{T}; X \leftarrow F_{a^i}(b);$
$j \xleftarrow{\$} \mathbb{T}; Y \leftarrow F_{a^j}(b);$	$j \xleftarrow{\$} \mathbb{T}; Y \leftarrow F_{a^j}(b);$
$Z \leftarrow F_{a^{i+j}}(b);$	$\ell \xleftarrow{\$} \mathbb{T}; Z \leftarrow F_{a^\ell}(b);$
$b \leftarrow \mathcal{A}(X, Y, Z);$	$b \leftarrow \mathcal{A}(X, Y, Z);$
return $b$ .	return $b$ .

Now define the *advantage* of  $\mathcal{A}$  violating the CSP-based decisional Diffie–Hellman assumption as

$$\mathbf{Adv}_{F,\mathcal{A}}^{csp-ddh} = |\Pr[\mathbf{Exp}_{F,\mathcal{A}}^{csp-ddh-real} = 1] - \Pr[\mathbf{Exp}_{F,\mathcal{A}}^{csp-ddh-rand} = 1]|. \quad (5.25)$$

That is, the CSP-DDH assumption states that the distributions  $(F_{a^i}(b), F_{a^j}(b), F_{a^{i+j}}(b))$  and  $(F_{a^i}(b), F_{a^j}(b), F_{a^\ell}(b))$  are computationally indistinguishable when  $i, j, \ell$  are drawn at random from  $\mathbb{T}$ .

We now move on to another assumption that will be used in our proposal. Here, the adversary is allowed to access a restricted CSP-DDH oracle  $\mathcal{O}_v(\cdot, \cdot)$ , which behaves as follows:

$$\mathcal{O}_v(X, U) = \begin{cases} 1, & \text{if } U = F_v(X); \\ 0, & \text{otherwise.} \end{cases} \quad (5.26)$$

That is, the oracle tells whether the second argument is conjugate to the first argument by  $v$ . This oracle can be treated as a restricted CSP-DDH oracle with one of its arguments being fixed as  $F_v(b)$ .

**Definition 5.4.9** (CSP-based Strong Diffie–Hellman Problem: CSP-SDH). Let  $F$  be a Conj-LD system over a noncommutative monoid  $G$  and let  $\mathcal{A}$  be an adversary. Given

$a \in G^{-1}$  and  $b \in G$ , consider the following experiment

Experiment  $\mathbf{Exp}_{F,\mathcal{A}}^{csp-sdh}$

$$i \xleftarrow{\$} \mathbb{T}; X \leftarrow F_{a^i}(b);$$

$$j \xleftarrow{\$} \mathbb{T}; Y \leftarrow F_{a^j}(b);$$

$$\mathcal{O}_v(X, U) \stackrel{def}{=} (U = F_v(X));$$

$$Z \leftarrow \mathcal{A}^{\mathcal{O}_v(\cdot, \cdot)}(X, Y);$$

if  $Z = F_{a^{i+j}}(b)$  then  $b \leftarrow 1$  else  $b \leftarrow 0$ ;

return  $b$ .

Now define the *advantage* of  $\mathcal{A}$  violating the CSP-based strong Diffie–Hellman assumption as

$$\mathbf{Adv}_{F,\mathcal{A}}^{csp-sdh} = \Pr[\mathbf{Exp}_{F,\mathcal{A}}^{csp-sdh} = 1]. \quad (5.27)$$

The intuition behind the CSP-SDH assumption is that the restricted CSP-DDH oracle is useless because the adversary already “knows” the answer to almost any query it will ask. Similar to [77], it is also worth mentioning that the CSP-SDH assumption is different from (and weaker than) the so-called *gap CSP-CDH assumption* where an adversary gets access to a *full* CSP-DDH decision oracle.

*Remark 5.4.10.* At present, it is unclear whether the CSP-DDH problem is indeed hard or not. Intuitively, we cannot solve the CSP-DDH problem without solving the CSP problem if  $G$  is a generic monoid. In fact, the CSP problem and the CSP-DDH problem of Conj-LD systems over a noncommutative monoids are counterparts of the DLP problem and the DDH problem over finite fields, respectively. According to [212], we know that the DLP problem and the DDH problem are polynomially equivalent in a generic cyclic group. Therefore, we speculate that the CSP problem and the CSP-DDH problem in a generic noncommutative monoid are polynomially equivalent.

### 5.4.3 Cryptosystems from Conj-LD Systems

In this section, we always assume that  $F$  is a Conj-LD system defined over a non-commutative monoid  $G$ , and  $a \in G^{-1}$  and  $b \in G$ .

As a warmup step, let us at first describe a Diffie–Hellman-like key agreement protocol by using the CSP-DDH assumption over a noncommutative monoid  $G$ . Assume that Alice and Bob want to negotiate a common session key. Then, Alice (resp. Bob) picks  $s \in \mathbb{T}$  (resp.  $t \in \mathbb{T}$ ). Then, send  $F_{a^s}(b)$  (resp.  $F_{a^t}(b)$ ) to Bob (resp. Alice). Finally, both of them can extract  $F_{a^{s+t}}(b)$ , by which a session key can be computed as

$$K_{\text{session}} = \text{Kdf}(F_{a^{s+t}}(b)), \quad (5.28)$$

where  $\text{Kdf}(\cdot)$  is a key derivation function, such as KDF1 defined in IEEE Std 1363-2000.

The above protocol<sup>10</sup> immediately implies the following ElGamal-like construction, denoted by CSP-ElG.

**Protocol 5.4.11.** CSP-Based ElGamal-Like Encryption Scheme (CSP-ElG):

- **Key-generation.** Suppose that  $k$  is the security parameter,  $\mathcal{M} = \{0, 1\}^k$  is the message space, and  $\mathcal{C} = \mathcal{K}_{[a,b]} \times \mathcal{M}$  is the ciphertext space. In addition, we need a cryptographic hash function  $H : \mathcal{K}_{[a,b]} \rightarrow \mathcal{M}$ . A user at first picks  $s \in \mathbb{T}$ , and then publishes  $pk = F_{a^s}(b)$  as his public key, and keeps his secret key  $sk = s$ .
- **Encryption.** Given the public-key  $pk \in \mathcal{K}_{[a,b]}$  and a message  $m \in \mathcal{M}$ , one picks  $t \in \mathbb{T}$ , and then constructs a ciphertext as follows:

$$c = (F_{a^t}(b), m \oplus H(F_{a^t}(pk))) \quad (5.29)$$

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<sup>10</sup>This protocol cannot resist the so-called man-in-middle attack. But in this section we don't pay attention to how to remedy it.



- **Decryption.** Given the secret key  $s \in \mathbb{T}$  and a ciphertext  $c = (c_1, c_2) \in \mathcal{K}_{[a,b]} \times \mathcal{M}$ , one can extract a plaintext as follows:

$$m = c_2 \oplus H(F_{a^s}(c_1)) \quad (5.30)$$

The consistency of the CSP-ElG scheme directly comes from the power law of the Conj-LD system. The security of the CSP-ElG scheme is captured by the following theorem.

**Theorem 5.4.12** (IND-CPA of CSP-ElG). *Based on the CSP-DDH assumption, we claim that the ciphertexts of the encryption scheme CSP-ElG are indistinguishable under chosen plaintext attacks in the standard model.*

*Proof.* Assume that the CSP-DDH assumption holds for the underlying noncommutative monoid  $G$ . We will prove by contradiction that CSP-ElG is IND-CPA. Suppose that CSP-ElG is not IND-CPA, and let  $\mathcal{A}$  be an algorithm which, on the system parameters  $a \in G^{-1}$ ,  $b \in G$  and a random public key  $pk \in \mathcal{K}_{[a,b]}$ , has probability non-negligibly greater than  $1/2$  of distinguishing random encryptions  $Enc(m_0)$  and  $Enc(m_1)$  of two messages  $m_0, m_1$  of its choice. Let  $Z = (F_{a^s}(b), F_{a^t}(b), F_{a^u}(b)) \in \mathcal{K}_{[a,b]}^3$  be either a random CSP-DDH triple or a random triple, with equal probability. We will produce an algorithm  $\mathcal{B}$  which can distinguish between the two cases, using  $\mathcal{A}$  as an oracle, with high probability.

The algorithm  $\mathcal{B}$  who interacts with the algorithm  $\mathcal{A}$  can be defined as follows:

- $\mathcal{B}$  at first picks a random integer  $v \in \mathbb{T}$  and sets the public-key as  $pk = F_{a^{s+v}}(b) \in \mathcal{K}_{[a,b]}$ , which is then sent to  $\mathcal{A}$ .
- Upon receiving the public-key  $pk$ ,  $\mathcal{A}$  selects two messages  $m_0, m_1 \in \mathcal{M}$  with equal length as the challenge messages, which are then sent to  $\mathcal{B}$ .
- Upon receiving the challenge messages pair  $(m_0, m_1)$ ,  $\mathcal{B}$  randomly picks another

integer  $w \in \mathbb{T}$ , flips a coin  $\beta \in \{0, 1\}$  and then replies  $\mathcal{A}$  with the challenge ciphertext as follows:

$$c_\beta^* = (F_{a^{t+w}}(b), m_\beta \oplus H(F_{a^{u+v+w}}(b))).$$

- Upon receiving the challenge ciphertext  $c_\beta^*$ ,  $\mathcal{A}$  replies  $\mathcal{B}$  with  $\hat{\beta} \in \{0, 1\}$ , i.e.,  $\mathcal{A}$ 's guess on  $\beta$ .
- Upon receiving  $\hat{\beta} \in \{0, 1\}$ , i.e.,  $\mathcal{A}$ 's guess on  $\beta$ ,  $\mathcal{B}$  checks whether  $\mathcal{A}$ 's guess is correct, i.e., whether  $\hat{\beta} = \beta$  holds.
- By repeating the above process with several different choices of random integers  $v, w$ , the algorithm  $\mathcal{B}$  can determine with high probability whether or not  $\mathcal{A}$  can determine the value of  $\beta$ . Based on the detailed analysis given below  $\mathcal{B}$  can, in this way, determine whether or not  $Z$  is a CSP-DDH triple, thus violating the CSP-DDH assumption over  $G$ .

In detail, there are two cases to be taken into consideration during a single execution of the above interactive process between  $\mathcal{B}$  and  $\mathcal{A}$ .

- Suppose that  $u = s + t$ . Then  $u + v + w = (s + v) + (t + w)$ , so  $Z$  is a CSP-DDH triple in  $\mathcal{K}_{[a,b]}^3$ . Moreover, all possible CSP-DDH triples w.r.t.  $(a, b)$  are equally likely to occur as  $Z$ , since  $v$  and  $w$  are random. Hence,  $c_\beta^*$  is a valid random encryption of  $m_\beta$  because  $F_{a^{t+w}}(b)$  is random in  $\mathcal{K}_{[a,b]}$ . Under these conditions, the algorithm  $\mathcal{A}$ , by hypothesis, will succeed in outputting  $\beta$  with probability exceeding  $1/2$  by a non-negligible quantity.
- Suppose that  $u$  is random. Then  $Z$  is a random triple in  $\mathcal{K}_{[a,b]}^3$ , and all possible triples belonging to  $\mathcal{K}_{[a,b]}^3$  occur with equal probability. In this situation, the probability distribution of  $c_0^*$  is identical to that of  $c_1^*$ , w.r.t. all possible random

choices of  $v$  and  $w$ . It follows that the algorithm  $\mathcal{A}$  cannot exhibit different behavior for  $\beta = 0$  and  $\beta = 1$ . Note that we can arrive at this conclusion even though the expression  $c_\beta^*$  is an invalid encryption of  $m_\beta$  — that is, even though there is no information about how  $\mathcal{A}$  behaves on invalid inputs, it is certain that  $\mathcal{A}$  cannot behave differently depending on the value of  $\beta$ .

The above analysis shows that if  $Z$  is a CSP-DDH triple then  $\mathcal{A}$  with non-negligible probability exhibits different behavior depending on whether  $\beta = 0$  or  $\beta = 1$ , whereas  $\mathcal{A}$  must behave identically regardless of the value of  $\beta$  if  $Z$  is not a CSP-DDH triple.  $\square$

Based on the above scheme, it is not difficult to derive a CCA secure encryption scheme by employing the Fujisaki-Okamoto transformation. Here, we would like to give two different extensions that are enlightened by [1, 77, 91, 177]. The first extension is called the hashed ElGamal variant that is IND-CCA secure in the random oracle model, and the second is called the Cramer-Shoup-like variant that is IND-CCA secure in the standard model.

**Protocol 5.4.13.** CSP-hElG.

- **Key-generation.** Suppose that  $k$  is the security parameter,  $\mathcal{M} = \{0, 1\}^k$  is the message space, and  $\mathcal{C} = \mathcal{K}_{[a,b]} \times \mathcal{M}$  is the ciphertext space. In addition, we need a symmetric cipher  $\Pi = (E, D)$  with the key space  $\mathcal{K}$  and a hash function  $H : \mathcal{K}_{[a,b]}^2 \rightarrow \mathcal{K}$ . A user at first picks an integer  $s \in \mathbb{T}$ , and then publishes  $pk = F_{a^s}(b)$  as his public key, and keeps his secret key  $sk = s$ .
- **Encryption.** Given the public-key  $pk \in \mathcal{K}_{[a,b]}$  and a message  $m \in \mathcal{M}$ , one picks an integer  $t \in \mathbb{T}$ , and then constructs a ciphertext  $c = (c_1, c_2)$  as follows:

$$c_1 := F_{a^t}(b), T := F_{a^t}(pk), K := H(c_1, T), c_2 := E_K(m) \quad (5.31)$$

- **Decryption.** Given the secret key  $s \in \mathbb{T}$  and a ciphertext  $c = (c_1, c_2) \in \mathcal{K}_{[a,b]} \times$

$\mathcal{M}$ , one can extract a plaintext as follows:

$$Z := F_{a^s}(c_1), K := H(c_1, Z), m := D_K(c_2) \quad (5.32)$$

The consistency of the CSP-hElG scheme can also be easily verified according to the power law of the Conj-LD system. The security of the CSP-hElG scheme is formulated by the following theorem.

**Theorem 5.4.14** (IND-CCA of CSP-hElG). *If  $H$  is modeled as a random oracle, and the underlying symmetric cipher  $\Pi$  is secure against chosen ciphertext attacks, then the hashed ElGamal encryption scheme CSP-hElG is secure against chosen ciphertext attacks under the strong CSP-CDH assumption.*

*Proof.* See Theorem 2 in [1] and Theorem 1 in [177], as well as the improvement of DHIES given in Section 5.1 of [177].

**Protocol 5.4.15.** CSP-CS.

- **Key-generation.** Suppose that  $k$  is the security parameter,  $\mathcal{M} = \{0, 1\}^k$  is the message space, and  $\mathcal{C} = \mathcal{K}_{[a,b]}^2 \times \widetilde{\mathcal{K}_{[a,b]}} \times \mathcal{M}$  is the ciphertext space, where  $\widetilde{\mathcal{K}_{[a,b]}}$  is defined as

$$\widetilde{\mathcal{K}_{[a,b]}} \triangleq \{F_{a^i}(F_{a^{j_1}}(b)F_{a^{j_2}}(b)) : i, j_1, j_2 \in \mathbb{T}\}.$$

In addition, we need a symmetric cipher  $\Pi = (E, D)$  with the key space  $\mathcal{K}$  and two hash functions  $H : \mathcal{K}_{[a,b]}^2 \rightarrow \mathbb{T}$  and  $H_1 : \mathcal{K}_{[a,b]} \rightarrow \mathcal{K}$ . A user at first picks  $x_1, x_2, x_3, x_4 \in \mathbb{T}$ , and then publishes  $pk = (X_1, X_2, X_3, X_4)$  as his public key, and keeps his secret key  $sk = (x_1, x_2, x_3, x_4)$ , where  $X_i = F_{a^{x_i}}(b)$  for  $i = 1, 2, 3, 4$ .

- **Encryption.** Given the public-key  $pk = (X_1, X_2, X_3, X_4) \in \mathcal{K}_{[a,b]}^4$  and a message  $m \in \mathcal{M}$ , one picks  $t \in \mathbb{T}$ , and then constructs a ciphertext  $c = (c_1, c_2, c_3, c_4)$  as follows:

$$c_1 := F_{a^t}(b), \quad c_2 := F_{a^t}(X_1), \quad c_3 := F_{a^t}(F_{a^h}(X_2)X_3), \quad c_4 := E_K(m),$$

where  $h := H(c_1, c_2)$  and  $K := H_1(F_{a^t}(X_4))$ .

- **Decryption.** Given the secret key  $(x_1, x_2, x_3, x_4) \in \mathbb{T}^4$  and a ciphertext  $c = (c_1, c_2, c_3, c_4) \in \mathcal{K}_{[a,b]}^2 \times \widetilde{\mathcal{K}_{[a,b]}} \times \mathcal{M}$ , one at first computes  $h := H(c_1, c_2)$  and then tests if the following two equalities hold:

$$F_{a^{x_1}}(c_1) = c_2 \text{ and } F_{a^{h+x_2}}(c_1)F_{a^{x_3}}(c_1) = c_3.$$

If not, output  $\perp$ . Otherwise, compute  $K := H_1(F_{a^{x_4}}(c_1))$  and output  $m := D_K(c_4)$ .

The consistency and the security of the CSP-CS scheme are formulated by the following theorems, respectively.

**Theorem 5.4.16** (Consistency of CSP-CS). *The CSP-CS scheme is consistent.*

*Proof.* Suppose that  $c = (c_1, c_2, c_3, c_4)$  is a well-formed ciphertext on the message  $m$ . Then, there exists some integer  $t \in \mathbb{T}$  so that

$$c_1 := F_{a^t}(b), \quad c_2 := F_{a^t}(X_1), \quad c_3 := F_{a^t}(F_{a^h}(X_2)X_3), \quad c_4 := E_K(m),$$

where  $h := H(c_1, c_2)$  and  $K := H_1(F_{a^t}(X_4))$ . Since  $X_i = F_{a^{x_i}}(b)$  for  $i =$

1, 2, 3, 4, we have

$$\begin{aligned}
 K &= H_1(F_{a^t}(F_{a^{x_4}}(b))) = H_1(F_{a^{t+x_4}}(b)) \\
 &= H_1(F_{a^{x_4}}(F_{a^t}(b))) = H_1(F_{a^{x_4}}(c_1)), \\
 c_2 &= F_{a^t}(F_{a^{x_1}}(b)) = F_{a^{x_1+t}}(b) \\
 &= F_{a^{x_1}}(F_{a^t}(b)) = F_{a^{x_1}}(c_1),
 \end{aligned}$$

$$\begin{aligned}
 c_3 &= F_{a^t}(F_{a^h}(F_{a^{x_2}}(b))F_{a^{x_3}}(b)) = F_{a^t}(F_{a^{h+x_2}}(b)F_{a^{x_3}}(b)) \\
 &= F_{a^t}(F_{a^{h+x_2}}(b))F_{a^t}(F_{a^{x_3}}(b)) = F_{a^{t+h+x_2}}(b)F_{a^{t+x_3}}(b) \\
 &= F_{a^{h+x_2}}(F_{a^t}(b))F_{a^{x_3}}(F_{a^t}(b)) = F_{a^{h+x_2}}(c_1)F_{a^{x_3}}(c_1).
 \end{aligned}$$

That is, the ciphertext  $c$  can stand the validation and the value of  $K$  used in encryption  $c_4 = E_K(m)$  is exactly the value of  $K$  used in decryption  $m = D_K(c_4)$ . Since  $\Pi = (E, D)$  is symmetric, the output  $m$  of the decryption is exactly the input  $m$  of the encryption. Therefore, the CSP-CS scheme is consistent.

**Theorem 5.4.17** (IND-CCA of CSP-CS). *Suppose  $H$  is a target collision resistant hash function. Further, suppose the CSP-DDH assumption holds, and the symmetric cipher  $\Pi = (E, D)$  is secure against chosen ciphertext attack. Then CSP-CS is secure against chosen ciphertext attack.*

*Proof.* See Theorem 13 in [77].

#### 5.4.4 Security and Efficiency Issues on $F_{a^t}(b)$

For computing  $F_{a^t}(b)$ , we should at first compute  $a^t$ , and then plus one inversion and two multiplications in the underlying noncommutative monoid  $G$ . When  $t$  is large, say several hundreds of digits, rather than to multiply  $a$  for  $t$  times, a similar “successive

doubling” method should be employed, and thus a factor of  $\log t$  would be taken into consideration in performance evaluations. At present, it is enough to set  $t$  as an integer with 160 bits to resist exhaustive attacks.

It is necessary to assume that the basic monoid operations (i.e., multiplication and inversion) can be finished efficiently. This assumption implies that the lengths of the representations of all elements in  $G$ , including  $a$ ,  $b$ ,  $a^t$ , and  $F_{a^t}(b)$ , should be polynomial in the system security parameters, since the results have to be provided in bits by using classical computers.

Moreover, we require that there is a *secure efficient canonical form* for representing elements in  $G$ . This means that

- (C-1) By using this form, the representation of an element in  $G$  is unique. Otherwise, the proposed schemes cannot work.
- (C-2) The transformation from an element in  $G$  to its canonical form can be finished efficiently. Otherwise, the proposed schemes are impractical.
- (C-3) By using this form, the length of an element  $F_{a^t}(b)$  does not reveal any information about  $a^t$ . Otherwise, the developed assumptions could suffer from the so-called length-based attacks.

## 5.5 Improved Key Exchange over Thompson’s Group

In 2005, Shpilrain and Ushakov proposed a key exchange protocol, referred to as SU05, based on the intractability of *decomposition problem* on Thompson’s group [259]. They claimed that the decomposition problem has the same difficulty level with *conjugacy search problem* [6, 175], and the advantage of the proposed protocol is that it is invulnerable to *length-based attack* according to their simulation results. In 2007, Ruinskiy et al. [235] proposed an improved length-based attack, referred to

as RST07, which does indeed threaten the security of SU05. They used five methods, “using memory,” “avoiding repetitions,” “looking ahead,” “automorphism attacks,” and “alternative solutions,” to improve the basic length-based attack and tested their success rates of breaking SU05. From their simulation results, the best success rate is over 80%. By investigating the properties of SU05 and the basic length-based attack, as well as the computational complexity of RST07, we [80] propose an improved key exchange protocol over Thompson's Group. It can resist known length-based attacks, including RST07 and other similar ones.

### 5.5.1 Thompson's Group and Decomposition Problem

Thompson's group  $F$  is well known in many areas of mathematics, including algebra, geometry, and analysis. This group is infinite and noncommutative. It is important that Thompson's group has the following nice presentation in terms of generators and relations:

$$F = \langle x_0, x_1, x_2, \dots \mid x_i^{-1} x_k x_i = x_{k+1} (k > i) \rangle \quad (5.33)$$

This presentation is infinite. There are also finite presentations of this group, for example,

$$F = \langle x_0, x_1, x_2, x_3, x_4 \mid x_i^{-1} x_k x_i = x_{k+1} (k > i, k < 4) \rangle$$

For the infinite presentation, any  $w \in F$  can be represented as a unique normal form

$$x_{i_1} \cdots x_{i_s} x_{j_t}^{-1} \cdots x_{j_1}^{-1} \quad (5.34)$$

so that

- (NF1)  $i_1 \leq \cdots \leq i_s$  and  $j_1 \leq \cdots \leq j_t$
- (NF2) if both  $x_i$  and  $x_i^{-1}$  occur, then either  $x_{i+1}$  or  $x_{i+1}^{-1}$  occurs, too.



The time complexity of reducing a word of length  $n$  to the normal form in Thompson's group is  $O(|n| \log |n|)$ , which is very efficient. Since each word has a unique normal form, we use “word length” to represent “length of the word's normal form” in the rest of this section.

Recall the decomposition problem over braid group (See Section 5.2.2), we define the decomposition problem over Thompson's group  $F$  as follows: Given a subset  $A \subseteq F$  and a pair of element  $(w, w') \in F^2$ , find two elements  $x, y \in A$  so that  $w' = x \cdot w \cdot y$ . In fact, the conjugacy search problem is a special case of the decomposition problem in the sense of taking  $x = y^{-1}$ .

### 5.5.2 Analysis of SU05 Protocol

Based on the intractability assumption of the decomposition problem over arbitrary (semi) group  $G$ , achieving key agreement is quite straightforward: given two subsets  $A, B \subseteq G$  so that  $ab = ba$  for any  $a \in A, b \in B$ , and given a public element  $w \in G$ . Alice selects private  $a_1 \in A, b_1 \in B$  and sends the element  $a_1 w b_1$  to Bob. Similarly, Bob selects private  $b_2 \in B, a_2 \in A$  and sends the element  $b_2 w a_2$  to Alice. Then Alice computes  $K_{\text{Alice}} = a_1 b_2 w a_2 b_1$  and Bob computes  $K_{\text{Bob}} = b_2 a_1 w b_1 a_2$ . Since  $a_i b_j = b_j a_i$  ( $i, j = 1, 2$ ) in  $G$ , one has  $K_{\text{Alice}} = K_{\text{Bob}} = K$ , which can be used as Alice's and Bob's common secret key.

The non-triviality of the SU05 protocol lies in instantiating with the Thompson's group  $F$  given by its standard infinite presentation (5.33) and the method for defining the suite subset  $A, B \subseteq F$ , as well as the method for sampling related elements in  $F$ .

Suppose that  $s \in N^+$  is the first system parameter, which is relevant to the size of following *generating groups*. Let  $S_A = \{x_0 x_1^{-1}, x_0 x_2^{-1}, \dots, x_0 x_s^{-1}\}$ ,  $S_B = \{x_{s+1}, x_{s+2}, \dots, x_{2s}\}$ ,  $S_W = \{x_1, x_2, \dots, x_{s+2}\}$  be the generating groups, and define the groups generated by  $S_A^{\pm 1}, S_B^{\pm 1}, S_W^{\pm 1}$  as  $A_s, B_s$  and  $W$ , respectively. Obviously,  $A_s$  and  $B_s$  are subgroups of  $F$  and according to [259], for  $\forall a \in A_s$  and  $\forall b \in B_s$ ,

$ab = ba$ .

Another system parameter of the protocol,  $L \in N^+$ , is chosen to control the word length. SU05 uses the same method to randomly generate  $a_1, a_2 \in A_s, b_1, b_2 \in B_s$  and  $w \in W$ . We take  $a_1$  as an example to analyze the generating process.

The generating process begins with an empty word  $e$ . Let  $u_0 = e$ . Firstly, randomly choose an element  $r_1$  from  $S_A^{\pm 1}$ , and multiply  $u_0$  on the right by  $r_1$  and get  $u_1$ . For example, if  $r_1 = x_0 x_1^{-1}$ , then  $u_1 = u_0 r_1 = x_0 x_1^{-1}$ . Repeat the previous step until the word length reaches  $L$ . Then we have generated a random word  $a_1 \in A_s$  of length  $L$ .

According to analysis of the generating process of  $a_1$ , we find that  $a_1$  can always be generated within finite steps from an empty word to a word of length  $L$ , due to one property of Thompson's group: with very high probability, product of two words has larger length than any one of the two words. Thus after each multiplication, the word length increases with high probability. As long as the generators are randomly chosen, the word length will finally reach  $L$  within finite steps. Suppose the expectation of word length increment after each step is  $E(\Delta l)$ , then on average after  $L/E(\Delta l)$  steps, a word of length  $L$  can be generated. Here  $E(\Delta l)$  is a constant which only relates to subset  $S_A^{\pm 1}$ , so the generating time of  $a_1$  is linear in  $L$ .

Although the generating process is efficient due to the property of Thompson's group we mentioned above, the protocol is also insecure under the length-based attacks due to the same reason.

### 5.5.3 Analysis of RST07 Attack

At first, let us review the process of the basic length-based attack and analyze the capability of SU05 for resisting the basic length-based attack. Let  $G$  be a subgroup of Thompson's group  $F$ , generated by the generators in  $S_G = \{g_1^{\pm 1}, \dots, g_k^{\pm 1}\}$ . For any unknown  $y, w \in G$ , given the product  $z = yw$  and a black box **BOX**, where for any  $w' \in G$ , if  $w' = w$  holds,  $\mathbf{BOX}(w') = 1$ , otherwise  $\mathbf{BOX}(w') = 0$ . The purpose of

length-based attack is to output the exact  $w$ .

Represent  $y$  as the product of generators, i.e.,  $y = y_1 y_2 \cdots y_n$ , where  $y_i \in S_G$ . Then, we have  $z = y_1 y_2 \cdots y_n w$ . Let  $\ell(z)$  be the word length of  $z$ . Since in Thompson's group, "with very high probability, product of two words has larger length than any one of the two words," the following inequality also holds with very high probability

$$\ell(y_1^{-1} z) < \ell(z) < \ell(y_j z), \quad (5.35)$$

where  $y_j \in S_G$  and  $y_j \neq y_1^{-1}$ . It indicates that if we multiply each of the  $2k$  generators on the left of  $z$ , except  $y_1^{-1}$ , other  $2k - 1$  generators will make the length of product increase with very high probability.

Based on this observation, a basic length-based attack algorithm can be formalized as follows:

---

**Algorithm 1 Basic Length-based Attack**

---

**Procedure** Basic Length-based Attack

**Input:** Product  $z = yw$ , black box **BOX**

**Output:**  $w$  or  $\perp$

```

1:   $z \leftarrow yw$ 
2:  while running time does not exceed limitation do
3:      choose  $h \in S_G$  so that  $\ell(h^{-1}z) \leq \ell(g^{-1}z)$  for all  $g \in S_G$ 
4:       $z \leftarrow h^{-1}z$ 
5:      if BOX( $z$ ) = 1 then
6:          return  $z$ 
7:      end if
8:  end while
9:  return  $\perp$ 
```

---

Note that the algorithm needs a limitation of running time in order to stop looping if it keeps failing within a long time. As we can see, the algorithm is based on the assumption that the generator  $h$  which minimizes  $\ell(h^{-1}z)$  is the right decomposition.

We take the following example to analyze the attack algorithm. Assume  $k = 3$ , i.e.,  $S_G = \{g_1^{\pm 1}, g_2^{\pm 1}, g_3^{\pm 1}\}$ . Let  $n = 3$  and  $y = g_1 g_3 g_2^{-1}$ . The process of generating  $y$  can be seen as operation on a  $2k$ -ary tree, as shown in Fig 5.6(a). It begins with an empty word  $e$  and one of the  $2k(k = 3)$  child-nodes,  $y_1$ , which is randomly chosen. Then, repeat the previous step until some  $y_n$  is chosen so that the length of  $y = y_1 y_2 \cdots y_n$  is satisfied. Since the  $2k$ -ary tree has  $(2k)^n$  leaf-nodes at the  $n$ th level, where each leaf-node represents a possible value of  $y$ , it is impossible to find  $y$  with brute force search.

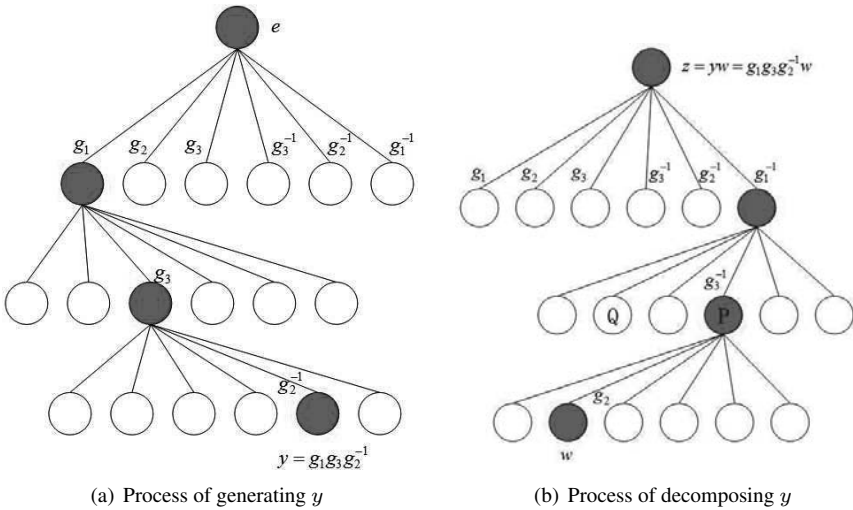


Figure 5.6: An example for analyzing basic length-based attack

The process of attack is reverse for instance searching on a  $2k$ -ary tree. Dark nodes are the targets of attack which forms the optimal search path (as shown in Figure 5.6(b)). The best result of an attack algorithm is to find this path which indicates decomposition of  $y$ . However, mostly there are more than one node to choose in step 5 of algorithm ??.

For example, in Fig 5.6(b), during the process of searching in the 2nd level, if there are two child-nodes P, Q with equal length and the algorithm chooses node Q wrongly, then the algorithm has to search all the child-nodes of Q which has smaller length than P before turning back to P. It makes the algorithm fall into an exponential search in the worst case. If there are on average two candidate nodes in each level, then the time complexity of attack algorithm is  $\mathcal{O}(2^n)$ . This should be the main reason why basic length-based attack cannot defeat SU05 [259].

Next, let us analyze the performance of RST07, i.e., the improved length-based attack proposed by Ruinskiy et al. [235]. RST07 includes the following five improvements. We will analyze them one by one with the purpose of finding out methods for resisting this improved attack.

- Using memory.** Using memory increases width of search. By using memory, the algorithm computes  $2km$  child-nodes from  $M$  subtrees and chooses from them the shortest  $M$  ones as the roots of subtrees for the next loop. As a result, the search width is enhanced from 1 to  $M$ . By this, the algorithm becomes efficient, but it is still exponential. For example, if in any loop, the right node is not included in the  $M$  candidates, then the algorithm may degenerate to exponential. Suppose there exists a  $y_j$  so that  $\ell(y_j w_{j+1}) < \ell(w_{j+1})$ , where  $w_{j+1} = y_{j+1} \cdots y_n$ , say  $y_j$  left multiplies  $w_{j+1}$  makes its length decrease. In this situation, the right node may not be included in the  $M$  candidates, because the length of child-node, referred to as  $y_j$ , increases, but it happens to be correct. In fact, this is the most common situation in which the basic length-based attack fails, and the success of the attack is based on the core assumption that direction of length descent is correct. However, the above situation is just opposite. If we introduce more like situations, the success rate of the attack algorithm will obviously decrease.

- **Avoiding repetitions.** Avoiding repetitions is an improvement usually seen in search. Since two nodes in the search tree may have the same value, the algorithm maintains a hash table to record values of visited nodes. If the same value appears again, then the node will be canceled from the candidates list. Therefore, avoiding repetitions not only improves efficiency of the algorithm, but also prevents the algorithm from trapping into a closed loop. The extra time and space costs involved in the method are negligible.
- **Look-ahead.** Look-ahead increases depth of search. The original algorithm searches one level each time, while the method searches  $p$  level each time. The problem of look-ahead is obvious. On one hand, the time complexity it involves increases exponentially with  $p$ . Because the time cost of traversing a  $p$ -level subtree is  $(2k)^p$ ,  $p$  should not be large. On the other hand, since  $p$  is not large, if multiple right nodes are increased in length, even if the algorithm looks ahead for  $p$  steps, it still cannot find the right path. For example, in a  $p$ -level subtree, if there are  $p/2$  right nodes being increased in length, then the right node at level  $p$  is also increased in length, thus the algorithm will again fall into an exponential search. Simulation results of RST07 prove this claim, for the algorithm success rate increases slightly by using look-ahead.
- **Automorphism attacks.** From simulation results of RST07, the improvement of automorphism attack is not satisfied. It indicates that in Thompson's group, at least under their parameters selection, random automorphism functions are highly correlative, thus effect of this method is negligible.
- **Alternative solutions.** Alternative solution is not a general method, it is a specific solution to SU05. Although this improvement is very efficient, it does not change the search complexity. In fact, this method helps the algorithm to find more potential right nodes which vastly decreases search time, but it is based on

the condition that the search direction is right. If the algorithm enters a wrong subtree, it still has large possibility to fall into an exponential search.

In brief, the basic length-based attack frequently falls into exponential search, so its success rate is very low, while the most significant improvements of RST07 are using memory and look-ahead, because to some extent they are able to avoid exponential search. However, the cost of look-ahead is too large to make it effective. Avoiding repetition is usually seen in search algorithm, although it is effective, the algorithm complexity is not changed. Automorphism attacks are restricted by the highly related automorphism functions, thus it is also not effective. Alternative solution not only enhances verification efficiency but also decreases search volume, thus it obviously increases success rate. But it does not change search complexity from exponential to polynomial, and it is not a general method.

Therefore, if we want to improve the security of the SU05, we need to make the attack algorithm fall into exponential search. In order to test efficiency of RST07, we realized SU05 and RST07 and acquired our own experiment data. Through the data analysis we propose our improved key exchange protocol on Thompson's group.

#### 5.5.4 Tests and Improvements

We choose to realize the combination of "using memory + avoid repetitions + alternative solutions," which has the highest success rate in the original paper. The simulation results are shown in Table 5.2, where  $\eta_1$  is the success rate of RST07 recovering  $a_1$ ,  $M = 1024$ . We change  $s$  and  $L$  to analyze success rates under different parameters, and use them as reference of latter experiments. Each success rate is the average value of over 1000 test cases, and the failure condition is searching over 100,000 nodes.

Because of the possible differences between experiment equipments and between definitions of failure (how many nodes visited), our simulation results are different from those in the original paper. But we will continue to use our results as reference to

Table 5.2: Success rate of recovering  $a_1$ 

L	$\eta_1/\%$					
	$s = 3$	$s = 4$	$s = 5$	$s = 6$	$s = 7$	$s = 8$
64	93.5	82.6	67.3	50.1	32.5	11.6
128	34.5	23.1	17.2	8.9	3.1	1.2
256	11.3	6.2	3.7	1.2	0.3	0
320	8.8	4.2	1.5	0.6	0.1	0

analyze our improvements of SU05.

As shown in Table 5.2, the success rate of the attack algorithm decreases dramatically with the increment of  $s$  and  $L$ . Irrespective of the increase in base number or exponent, efficiency of the algorithm decreases rapidly. The phenomenon itself shows the algorithm is exponential. Obviously, increasing protocol parameters is one of the direct methods to resist exponential attacks.

Based on the analysis aforementioned, we find out that randomly generating  $a_1$  leads to nondecreasing growth of its length. As a result, when left multiplying  $a_1$  to  $a_1wb_1$  repeatedly, the length of product decreases gradually, and this situation just satisfies the success condition of length-based attacks. Thus we need to control the process of generating  $a_1$  to improve its robustness against the length-based attacks. Therefore, we propose an improved key exchange protocol over Thompson's group as follows:

1. Pick  $s \in [6, 10]$  and  $L \in \{512, 514, \dots, 640\}$ . Set the security parameters as  $N, T, R$ . Choose  $w \in W$  so that  $\ell(w) = L$ . Publish  $s, L, w, N, T, R$ .
2. Alice randomly chooses secret  $a_1 \in A_s, b_1 \in B_s$  so that  $\ell(a_1) = L, \ell(b_1) = L$ , and computes  $u_1 = a_1wb_1$ .
3. Suppose  $a_1 = r_1r_2 \cdots r_n, r_i \in S_A^\pm, i = 1, \dots, n, b_1 = t_1t_2 \cdots t_m, t_j \in S_B^\pm, j = 1, \dots, m$ . Alice verifies the following conditions.
  - (a)  $n \geq N$  (and  $m \geq N$ )
  - (b)  $\ell(r_T^{-1}r_{T-1}^{-1} \cdots r_1^{-1}u_1) \geq \ell(u_1)$  (and  $\ell(u_1t_m^{-1}t_{m-1}^{-1} \cdots t_{m-T+1}^{-1}) \geq \ell(u_1)$ )



- (c) After  $a_1$ 's first  $T$  factors are left multiplied to  $u_1$  ( $b_1$ 's first  $T$  factors are right multiplied to  $u_1$ ), the length of the resulting word  $r \geq R$ .

If above conditions are all satisfied, Alice sends  $u_1$  to Bob. Otherwise, go to step 2 to regenerate  $a_1, b_1$  and compute  $u_1$ .

4. Bob generates  $a_2, b_2$  and computes  $u_2$  similarly as Alice does in step 2 and 3.
5. Alice computes  $K_{\text{Alice}} = a_1 u_2 b_1$  after getting  $u_2$  from Bob.
6. Bob computes  $K_{\text{Bob}} = b_2 u_1 a_2$  after getting  $u_1$  from Alice.

There are two differences between the new and the original. One is that the system parameters  $s$  and  $L$  are increased. The other is that  $a_1, b_1, a_2, b_2$  should satisfy the following requirements.

We introduced three new parameters  $N, T, R$  in our protocol. Alice (and Bob) can decide them by herself (himself) to achieve different security levels. We take  $a_1$  as an example to explain the three parameters. (1) Let the number of  $a_1$ 's factors be greater than  $N$ . The more factors  $a_1$  has, the larger the depth of the search tree is, the less likely the attack algorithm succeeds. (2) Suppose that the length of  $u_1$  does not decrease when it is left multiplied by the inverses of the first  $T$  factors of  $a_1$ . It indicates that some of the first  $T$  factors increase the length of  $u_1$ . (3) Compute the maximum and minimum lengths in the process of left multiplying inverses of  $a_1$ 's first  $T$  factors to  $u_1$ , and define  $r$  as the difference of maximum and minimum lengths. Let  $r \geq R$ . A protocol with larger  $r$  is harder to be defeated by the attack algorithm.

We repeat the simulation with our protocol to show its advantages. We still choose the combination of "using memory + avoid repetitions + alternative solutions" as attack algorithm, in which  $M = 1024$ .

The purpose of our simulation is to show how the success rates of the attack algorithm are affected by the three new parameters. In order to reduce the affection of other parameters, we choose  $s \in (3, 4)$  and  $L \in \{64, 128, 256\}$ .

- **Given lower bound of factors' number (parameter  $N$ ).** If  $N$  is too large, the efficiency of generating  $a_1$  will be low. Thus we should choose a proper  $N$ . To test success rate of RST07 recovering  $a_1$ , we fix  $T = 0, R = 0$  and change  $s, L, N$  respectively, as shown in Table 5.3.

Table 5.3: Success rate of recovering  $a_1$  by RST07 with a restricted  $N$ 

	$N$	$\eta_2/\%$	
		$s = 3$	$s = 4$
$L = 64$	120	55.6	45.9
	140	52.8	41.6
	160	49.0	40.2
	180	45.7	37.8
$L = 128$	240	24.3	17.4
	280	22.0	15.5
	320	20.4	14.9
	360	18.1	13.6
$L = 256$	480	7.6	4.4
	560	6.9	3.8
	640	6.7	3.5
	720	6.1	3.2

Obviously, for different  $L$  we need to choose different range of  $N$ . Comparing Table 5.3 with Table 5.2, we find that the success rate decreases rapidly. For the same pair of  $s$  and  $L$ , increasing  $N$  decreases the success rate slightly. Thus, we just need to choose a proper  $N$ .

- **Ensure the number of factors increasing (parameter  $T$ ).**  $T$  is used to make sure that the length of the resulting word does not decrease when the inverses of the first  $T$  factors of  $a_1$  are left multiplied to  $u_1$ . Similarly, we should choose a proper  $T$  to guarantee the efficiency of generating  $a_1$ . We fix  $N = 0, R = 0$  and change  $s, L, T$  respectively, as shown in Table 5.4, to test success rate of RST07 recovering  $a_1$ .

From Table 5.4 we can see that this method greatly decreases success rates with an increment of  $T$ . Note that with the increment of  $T$ , the time of generating  $a_1$  is increasing. When  $T = 100$ , it spends about 30 minutes to acquire a legal  $a_1$ . This time is not acceptable. We suggest that  $T$  be less than 85.

Table 5.4: Success rate of recovering  $a_1$  by RST07 with a restricted  $T$ 

$T$	$\eta_3/\%$					
	$s = 3$			$s = 4$		
	$L = 64$	$L = 128$	$L = 256$	$L = 64$	$L = 128$	$L = 256$
40	33.1	12.6	6.3	31.6	10.7	4.7
50	25.3	9.5	4.2	22.9	7.5	2.2
60	12.5	6.7	2.3	10.2	4.2	1.0

- **Restriction for parameter  $R$ .**  $R$  is the difference between the maximum length and the minimum length. To test success rate of RST07 recovering  $a_1$ , we fix  $N = 0, T = 0$  and change  $s, L, R$  respectively.

Table 5.5: Success rate of recovering  $a_1$  by RST07 with a restricted  $R$ 

$R$	$\eta_4/\%$					
	$s = 3$			$s = 4$		
	$L = 64$	$L = 128$	$L = 256$	$L = 64$	$L = 128$	$L = 256$
6	36.4	18.6	8.7	32.5	13.7	4.9
8	32.1	16.9	8.0	29.8	11.5	4.5
10	27.5	13.9	6.8	26.6	9.5	3.9

As shown in Table 5.5, this method also effectively decreases success rates. The larger  $R$  is, the longer it takes for the attack algorithm to arrive at correct nodes. Similarly, the time of generating  $R$  increases when it becomes larger. So  $R$  should not be too large.

- **Combination of above improvements.** We take the three parameters into consideration simultaneously to design our last experiment. The choice of the parameters are shown in Table 5.6, where  $\eta_5$  is the success rate of RST07 recovering  $a_1$ .

Table 5.6: Success rate of recovering  $a_1$  by RST07 with restricted  $N, T, R$ 

$(N, T, R)$	$\eta_5/\%$					
	$s = 3$			$s = 4$		
	$L = 64$	$L = 128$	$L = 256$	$L = 64$	$L = 128$	$L = 256$
(120, 40, 6)	19.8	6.5	1.4	14.4	4.8	0.4
(140, 50, 8)	9.5	3.9	0.5	7.9	2.2	0.1
(160, 60, 10)	4.1	1.2	0.2	3.2	0.6	0

Comparing this result with Table 5.2, we see the improvements are effective.

Although the improved length-based attack RST07 is effective, it does still not change its complexity from exponential to polynomial, and its success is based on the randomly generating process. If we modify the process of generating words, the attack algorithm becomes inefficient. In addition, the idea of our improved protocol can be extended to resist general length-based attacks against the protocols over other algebraic structures. The core idea is that length-based attacks focus on the change of word lengths, so it can be used in all kinds of protocols over noncommutative groups.

## 5.6 Notes

In this chapter, we open the windows to noncommutative cryptography for those who are not familiar with this subject. Our introduction is far from complete. In particular, we did not mention other attacks against the aforementioned noncommutative cryptosystems. We think that a cryptographic scheme or protocol that opens a new door for cryptographic applications has definitely great significance, even if the original version is insecure in some ways. We also introduce some of our contributions on this subject, which are published at AsiaCCS 2007, Inscrypt 2010, and SCIENCE CHINA Information Sciences, etc. We refer to [80, 284–287] for more technical details. In addition, the readers are also suggested to read the survey papers [112] and the books [218].

As for braid-based cryptography, one of the most challenging problem is how to sample hard CSP instances. It turns out that for a braid-based cryptographic scheme to be secure, it is essential that keys are selected from a “very small” subset rather than from the whole group. Constructing these subsets for a particular cryptographic scheme is usually a very challenging problem [218].

As for  $\mathbb{Z}$ -modular method, there are at least three interesting and challenging directions that are to be developed. The first is to find more promising instantiations for various algebraic structures. The second is to design new cryptographic schemes

based on relevant cryptographic assumptions derived by  $\mathbb{Z}$ -modular method. The third is to extend the  $\mathbb{Z}$ -modular method from unitary polynomial to multivariate polynomial, and then bridge the  $\mathbb{Z}$ -modular method and the well-known multivariate cryptographic method.

Another interesting problem is to consider the braided Thompson group  $BV$  which was developed independently by Dehornoy and Brin. In geometry, an element in  $BV$  group can be viewed as the coupling of two elements in Thompson group  $F$  by using a braid (See Figure 5.7). According to Bux and Sonkin's suggestion,  $BV$  group does not admit non-trivial linear representation. Thus, from a cryptographic perspective, it possesses the advantages of braid group and Thompson's group simultaneously.

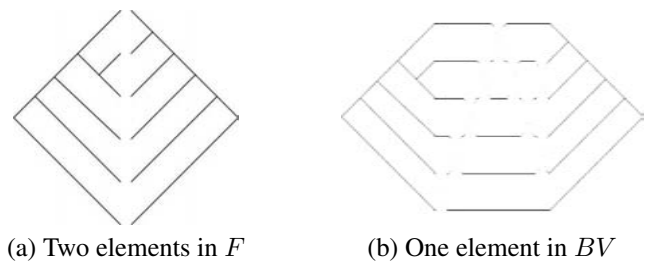


Figure 5.7: Geometric Illustration of Braided Thompson Group  $BV$

# Chapter 6

## Perspectives

In this chapter, we give some perspectives including some applications and some open problems of the new directions of modern cryptography. The applications cover all the wired and wireless communication networks, satellite communication networks, multicast/broadcast and TV networks, computer networks, all newly emerging networks, such as Internet of Things, Cloud Computing, Social Networks, Named Data Networks, and the fields of the digital right management protection, radio-frequency identification, proof of taken path, conditional access, biometric authentication, and so on.

### 6.1 On Proxy Re-Cryptography

As we see, more and more researchers are working on proxy re-cryptography, especially on its applications [245]. It is safe to say that applications drive forward the research on proxy re-cryptography.

Proxy re-cryptography is being applied to data sharing [104, 151, 168], group key management [81], cloud storage [197], vehicular ad hoc network (VANET) [269], mul-

roadcast [149], digital rights management (DRM) [163, 271], proof of taken path [11], anonymous routing [204], radio-frequency identification (RFID) [140, 170], certified mailing lists [171], and so on. These applications can be roughly classified into the following three kinds:

**Application 1: Inter-domain scenarios.** Proxy re-cryptography is originally designed for these scenarios, where the signature (ciphertext) is desired to be verifiable (decryptable) in many domains [11, 140, 163, 170, 197, 204, 269, 271]. When a signature (ciphertext) travels in a new domain, it can be transformed to the corresponding one in the new domain by using the transformation functionality of proxy re-cryptography. Once a signature (ciphertext) can travel across many domains as in its own domain, the problems of trust transfer, verifiability transfer, and decryption ability transfer can be solved easily, and the storage cost of different encryptions and signatures can be saved. Hence, the resulting system could become quite flexible and efficient.

**Application 2: Multi-receiver scenarios.** In these scenarios, one sender usually sends an encryption to many receivers via a third party or several third parties [81, 149, 168, 216]. In an ideal situation, no matter how many receivers exist and how many new receivers appear after the encryption, the computational cost of the encryption on the sender side is constant, and only the intended receivers can recover the plaintext. By using proxy re-encryption, the above demand can be easily met. In particular, the sender, the third party, and the receivers act as the delegator, proxy, and delegatee(s) in proxy re-encryption, respectively. When a new receiver appears, the sender only needs to send the corresponding re-encryption key to the proxy, and the proxy transforms the encryption to another encryption that can be decrypted by the new receiver.

**Application 3: Key revocation scenarios.** In these scenarios, the data is usually en-

encrypted under some keys and shared among a group of users via a server [81, 104, 151, 168]. Such scenarios require the backward security and the forward security, which essentially require a key revocation process. The backward security guarantees that the new user cannot access the old data, while the forward security prevents the user left from accessing new data. By using proxy re-encryption, once the data is re-encrypted by the server, it can be decrypted by the new decryption key (corresponding to the delegatee in proxy re-encryption), not the old decryption key (corresponding to the delegator in proxy re-encryption). Hence, the old decryption key is revoked.

Generally speaking, the normal proxy re-cryptographic schemes defined in Chapter 2 cannot be used in the above scenarios directly. For example, the cloud storage [197] requires the threshold proxy re-encryption where the role of the proxy is played by a group of sub-proxies. On the other hand, the current solutions based on proxy re-cryptography are not yet perfect. For instance, the solution to the proof of taken path [11] cannot efficiently prove the exact path the user took. It is because that we do not have any multiuse, unidirectional proxy re-cryptographic scheme with constant size.

Due to the above three situations in proxy re-cryptography, we can see a picture that there will be more new proxy re-cryptographic schemes designed for new applications, and more new applications based on proxy re-cryptography.

## 6.2 On Attribute-Based Cryptography

We present several application environments based on the research results in the attribute-based cryptography. As we know, attribute-based encryption could result in a perfect mergence of the encrypted storage and the access control [31, 55, 66, 188, 273, 280, 282, 306]. These applications can be roughly classified into the following three



kinds:

**Application 1: Multi and fuzzy targeted encryption.** Multi and fuzzy targeted encryption [227], is described as a new attribute-based encryption scenario. It means only one encrypted copy can be decrypted by different users with different private keys. Consider a scenario in which a large school periodically sends updated schedules to all the professors. Those files are encrypted under some access policies concerning a professor's gender ( $Male, \neg Male$ ), experience ( $(> 5 \text{ years}), \neg(> 5 \text{ years})$ ), department ( $Computer, \neg(Computer)$ ), and specialty ( $Network, \neg(Network)$ ). The school manager probably constructs an encrypted email with an access structure  $(Male) \wedge (> 5 \text{ years}) \wedge (Computer)$  to inform these professors the location of recent conferences. Suppose there is a male doctor, Jack, with seven years of work experience in a computer department. He is qualified to open the encrypted email since the attribute set of his secret key satisfies the access structure. However, a male professor in the Mathematics department is not qualified to open the email. As a direct result, this kind of encryption can be easily transformed to anonymous encryption schemes to preserve the receivers' privacy [32, 166, 189, 236]. We also want more applications on preserving privacy through this multi and fuzzy targeted encryption.

**Application 2: Trust authority application.** In the ABE schemes, the central authority can decrypt all the ciphertexts. We first remove the central authority by the threshold multi-authority [292]. Although no individual authority can decrypt all ciphertexts in the system, each authority can individually decrypt a part of ciphertexts whose access structures are satisfied by the attributes managed by the authority. We further solve this problem to achieve the trustfulness of the authority [201]. In addition, more scenarios can be applied to multi-authority ABE to decentralize the threats from the authority.

**Application 3: Conditional access system.** Ciphertext policy attribute-based encryption can be used in the conditional access system [31] of digital television [114, 305]. Different users can be labeled with different attributes according to their reservation or purchase. TV programs are encrypted by an access structure and sent to all users. Only the users whose attributes satisfy the access structure can watch the TV programs.

**Application 4: Location-based service.** Interval encryption can be used not only in the broadcast encryption schemes but also in the vehicle networks [15, 147, 150, 238] and other location-based service networks [157, 237]. The vehicle (equipped with the private key) can decrypt the ciphertext only if its private key belongs to the interval that is used in the encryption as a public key. No one else can decrypt without the related private keys. This kind of ciphertexts can be encrypted and stored at the service spots to construct networks for providing service. Interval encryption not only keeps the confidentiality of the message or service, but also provides a new solution to preserve the privacy in the vehicle networks [105].

**Application 5: Named data network application.** The emerging network, called data network (NDN) [310], focuses on the new Internet architecture that can capitalize on strengths and address weaknesses of the Internet's current host-based, point-to-point communication architecture in order to naturally accommodate patterns of communication. NDN has some innate privacy friendly features, such as lack of source and destination addresses in packets. In practice, this means that the adversary who eavesdrops on a link close to a content producer cannot immediately identify the consumer who expressed interest in that particular content. Moreover, two features of standard NDN routers, content caching and collapsing of redundant interests, reduce the utility of eavesdropping near a content producer since not all interests for the same content reach its producer [98]. However,

NDN provides no protection against an adversary that monitors local activities of a specific consumer. As most content names are expected to be semantically relevant to content itself, interests can leak a lot of information about the content they aim to retrieve. To mitigate this issue, “encrypted names” may be a solution. The producer encrypts the name. However, this simple approach does not provide names for routing or share the content in the requests. Attribute-based encryption could provide a perfect merge of the encrypted storage and the access control.

**Application 6: Selective attribute delegation.** Combined with the proxy re-encryption, attribute-based encryption could also be used in an application scenario of personal information system in a university [100]. In this system, there are some confidential records such as the grades of every student. This information is encrypted under the access policy “ $((AGE > 40) \wedge (Tenure))$ ”. The professors who are older than 40 and have a tenure position would receive the secret keys corresponding to “ $AGE > 40$ ”, “ $Tenure$ ”, and thus they are qualified to retrieve the confidential records. Nevertheless, when these professors are on vacation, it is necessary to find some trustworthy substitutes who can decrypt these records in time. Therefore, ABPRE allows a qualified professor to freely designate a proxy who can translate these encrypted records to those encrypted with a different access policy (such as administrators with at least 10 years of working experience “ $(Admin) \wedge (EXP > 10)$ ”). Hence, even if no qualified professor is available, some highly experienced administrator can open the confidential records with the help of the professor’s proxy [14, 154, 190, 239, 256].

**Application 7: Biometric authentication.** An important application of FIBS deals with biometric authentication [185, 220, 301]. A biometric authentication system is essentially a pattern-recognition system that recognizes a person based on a feature vector derived from a specific physiological or behavioral charac-

teristic that the person possesses. Since biometrics cannot be lost or forgotten like computer passwords, biometrics have the potential to offer higher security and more convenience for the users. However, two biometric feature scan results are not identical with high probability. As a result, it is not a good solution to directly choose biometric features as the password or authentication for comparison. FIBS system provides an attractive solution to biometric authentication using cryptographic techniques since two scan results have a certain (possibly high) proportion of overlaps for the same person while the proportion is low for the different persons. Using FIBS system [301], it is not necessary to store the biometric feature for comparison, which can save a lot of storage and achieve economic benefits.

### 6.3 On Batch Cryptography

Although the researches on batch cryptography started in 1989, there are still some open problems in the field. At the same time, applications based on the batch cryptography can be widely used. Let us conceive the perspective of batch cryptography. We present several application environments based on these research results on the batch cryptography.

**Application 1: Batch request process.** In the cloud computing environment [281, 292], the customer wants a handshake with the cloud server at the beginning of the communication. When the requests exceed the server's handling ability, the response will be delayed and the service would be influenced in the way similar to the deny of service attacks. It becomes worse when the customer uses encryption to send requests. Batch cryptography has resolved this problem. When the sender sends requests using the cloud server's public key, the cloud server can achieve batch key agreement to improve the efficiency and maintain the security as well.

It does not require the customers and the cloud server to do it in a point-to-point way. Here, we stress that it is an open problem how to apply the batch cryptography technique in every many-to-one communication environment to achieve efficiency and security.

**Application 2: Batch packet authentication.** In the wireless networks such as ad-hoc networks [307] and delay-tolerant networks [311,312], packet authentication is a critical security service that ensures authenticity and integrity of packets during multi-hop transmissions. Public key signatures, which have been suggested in existing packet security protocol specification, achieve packet authentication at the cost of increased computational overhead and transmission overhead, and a higher energy consumption, which is not desirable for energy-constrained ad hoc networks and DTNs. Batch authentication schemes, through aggregate signature and batch verification, could achieve efficient packet authentication and cut down the communication cost in the resource-constrained networks.

**Application 3: Batch network coding.** In the network coding scenarios, each packet is transmitted in the encrypted way by our batch RSA schemes. Each node in the networks can be equipped with the software or hardware, which can be designed as universal chips to parallelly decrypt the ciphertext in a fixed block. We can use the subset of solutions of Plus-Type Equations or Minus-Type Equations and use one modular operation to solve such a problem. These kinds of chips can also be used in the key exchange schemes for the multi-request scenarios in cloud computing.

## 6.4 On Noncommutative Cryptography

Although the beginning of the research on noncommutative cryptography can be traced back to 1984, this field is considerably young. As far as we know, the ear-

liest formal appearance of the term “noncommutative cryptography” is at the second international conference on symbolic computation and cryptography (SCC 2010) (See <http://scc2010.rhul.ac.uk/>). Informally, Myasnikov et al. used this term even earlier and the new upsurge of study on noncommutative cryptography owe a great deal of credits to their efforts. Now, let us conceive the perspective of noncommutative cryptography.

Recently, Fine et al. [112] gave a list of open problems related to group-based cryptography. Most of these open problems are relevant to noncommutative cryptography, too. Instead of enumerating each problem based on concrete noncommutative algebraic structures, we classify the open problems of noncommutative cryptography into the following categories:

**Open problem 1:** What is the most appropriate noncommutative platform (group, semigroup, ring, or other algebraic structures) for cryptography?

*In fact, this is one of the most common problems and it is very difficult to answer. At first, it seems that braid group might not be a good candidate, although it is probably the most “famous” platform for noncommutative cryptography. At present, we are even confronting the challenge of generating secure keys for braid-based cryptography, needless to say the recently published successful cryptanalysis on braid-based cryptographic schemes. Second, we believe that using special matrices in noncommutative cryptography has potential advantages. In particular, matrices have a ring structure, which means two binary operations rather than one can be used to “diffuse” transmissions. However, whenever matrices are used, there is always a danger of a linear algebra attack. A possible way to avoid this kind of attack might be using matrices over a rather peculiar semiring that does not admit*

*an embedding into a field. Third, as the well-known noncommutative group, the symmetric group  $S_n$  should be explored carefully. On one hand, the operations over  $S_n$  can be implemented efficiently and certain hidden subgroup problem over  $S_n$  can resist known quantum algorithm attacks; on the other hand, the CSP problem over  $S_n$  seems tractable on average — thus we have to pay attention to other hard problems over  $S_n$ .*

**Open problem 2:** How to estimate the hardness of the related cryptographic problems, especially when the problems are unsolvable or undecidable?

*The classical complexity theory always stops at the line when we conclude that the given computational/decisional problem is unsolvable/undecidable. However, this kind of conclusions are not sufficient for cryptographic applications. For example, Diophantine equations are unsolvable in general, but almost all published cryptographic schemes based on Diophantine equations are broken. This situation also happens for the knapsack problem which in the worst case is proved to be NP-hard. Therefore, to reduce the security of a cryptographic scheme to some underlying problem  $P$ , we should be confident on at least one of the following: (1)  $P$  is intractable in average case; (2)  $P$  is intractable in the worst case and we have an efficient method to sample many of the worst instances of  $P$ ; (3)  $P$  is intractable in the worst case and the reduction also worst-case, i.e., breaking of the scheme can lead to solving the arbitrary instance of  $P$ .*

**Open problem 3:** How to extend the theory of provable security so that it is suitable for noncommutative cryptography?

*The concept of indistinguishability plays a key role in the theory of*

*provable security. However, noncommutative cryptography always deals with finitely generated groups. Without proper sampling methods, it is hard to define a desired distribution, needless to say indistinguishability between distributions. In addition, random objects are very important in cryptography. Thus we need not only efficient methods to sample elements from a (probably infinite) space, but also well-established theory to estimate the randomness of the distributions obtained by given sampling methods. Fortunately, the probability theory on groups has been developed even before the birth of the public-key cryptography and now is available for analysis of probability measures and relevant characteristics of noncommutative cryptography.*

**Open problem 4:** Is it possible to bridge noncommutative cryptography with quantum cryptography?

*The starting point of noncommutative cryptography is to resist quantum attacks and we learn that quantum cryptography is unconditionally secure. Thus, it is meaningful to bridge between the two sides. On one hand, the development of quantum theory does indeed benefit from employing matrices theory and noncommutative multiplication [12]. In particular, each quantum gate can be expressed as a unitary matrix. The non-commutativity of matrix multiplication implies that quantum operations are essentially noncommutative too. Thus we can try to research quantum cryptography by employing the methods from noncommutative cryptography, such as the ideas behind braid-based cryptography. On the other hand, we have already developed some quantum algebraical objects and topological objects, such as quantum integers, quantum torus, quantum braids, and quan-*



*tum knots, etc. If the quantum versions of the corresponding classical cryptographic problems, say conjugator search problem over quantum braids, are intractable in the quantum Turing model, the quantum versions of the corresponding classical cryptographic schemes and protocols, say AAG protocol over quantum braids, might be secure against quantum attacks.*

**Open problem 5:** Is it possible to bridge noncommutative cryptography with biologic cryptography?

*The field of DNA computing, or, more generally, biomolecular computing, was established in Leonard M. Adleman's seminal paper [2]. DNA sequences are known to appear in the form of double helices in living cells, in which one DNA strand is hybridized to its complementary strand through a series of hydrogen bonds. DNA computing requires that the self-assembly of the oligonucleotide strands happen in such a way that hybridization should occur in a manner compatible with the goals of computation. There are multiple methods for building a computing device based on DNA, each with its own advantages and disadvantages. Among them, the concept of toehold exchange manifests a noncommutative algebraic system. In this system, an input DNA strand binds to a sticky end, or toehold, on another DNA molecule, which allows it to displace another strand segment from the molecule. This allows the creation of modular logic components such as AND, OR, and NOT gates and signal amplifiers, which can be linked into arbitrarily large computers. Clearly, this kind of binding operation is noncommutative. Therefore, the remaining problem is to probe intractable<sup>1</sup> problems based on toehold exchange system. Later, we can*

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<sup>1</sup>Here, the term *intractability* should be considered in DNA computational environment. In fact, from the

*design biologically cryptographic schemes.*

To conclude the exploration on the subject of noncommutative cryptography, we would like to quote paragraphs from the Fields Lecture that was given by Atiyah [12] at the World Mathematical Year 2000 Symposium:

*“A third theme is the shift from commutative to noncommutative. This is perhaps one of the most characteristic features of mathematics, particularly algebra, in the 20th century. The noncommutative aspect of algebra has been extremely prominent” . . . “All these are different ways or strands that form the basis of the introduction of noncommutative multiplication into algebra, which is the ‘bread and butter’ of 20th century algebraic machinery.”*

Is it possible that the introduction of noncommutative operations into cryptography would become the “bread and butter” of cryptographic machinery in the 21st century? We hope so.

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standpoint of computability theory, DNA computing does not provide any new capabilities. For example, if the space required for the solution of a problem grows exponentially with the size of the problem (i.e., EXPSPACE problems) on von Neumann machines, it still grows exponentially with the size of the problem on DNA machines. For very large EXPSPACE problems, the amount of DNA required is too large to be practical. (From Wikipedia, the free encyclopedia. [http://en.wikipedia.org/wiki/DNA\\_computing](http://en.wikipedia.org/wiki/DNA_computing))

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# Appendixes

## Appendix A: All 122 solutions of $\sum_{1 \leq i \leq 8} \frac{1}{x_i} + \prod_{1 \leq i \leq 8} \frac{1}{x_i} = 1$

- 1: (2,3,7,43,1807,3263443,10650056950807,113423713055421844361000443)
- 2: (2,3,7,43,1807,3263443,10650057792155,134811739261383753719)
- 3: (2,3,7,43,1807,3263443,10652778201539,41691378583707695)
- 4: (2,3,7,43,1807,3263443,10699597306267,2300171639909623)
- 5: (2,3,7,43,1807,3263447,2130014000915,22684798220467498090185211)
- 6: (2,3,7,43,1807,3263447,2130014387399,11739058070963394487)
- 7: (2,3,7,43,1807,3263479,288182779055,243811701792623)
- 8: (2,3,7,43,1807,3263483,260604226747,80249212735823)
- 9: (2,3,7,43,1807,3263495,200947673239,67137380077902268343)
- 10: (2,3,7,43,1807,3263495,200949404503,23316080984691959)
- 11: (2,3,7,43,1807,3263531,119666789791,8081907028348841339)
- 12: (2,3,7,43,1807,3263591,71480133827,761302020256877140089595)
- 13: (2,3,7,43,1807,3263779,31834629787,4396910340967)
- 14: (2,3,7,43,1807,3264187,14298637519,152316021000302785506427)
- 15: (2,3,7,43,1807,3316627,203509259,109643149191047)
- 16: (2,3,7,43,1807,3586039,36800447,2550097247)
- 17: (2,3,7,43,1811,655519,389313431,1507818475)
- 18: (2,3,7,43,1811,713899,7813583,2409102303622951)
- 19: (2,3,7,43,1811,793595,3722287,233296531681207)
- 20: (2,3,7,43,1817,298637,279594269,3859101523354821017)
- 21: (2,3,7,43,1819,252731,2134319143,6047845668256680791)
- 22: (2,3,7,43,1823,193667,637617223447,406555723635623909338363)
- 23: (2,3,7,43,1823,193667,637617223459,31273517203328870463055)
- 24: (2,3,7,43,1823,193675,4683210919,754794584867)
- 25: (2,3,7,43,1831,132347,231679879,1197240789041771)
- 26: (2,3,7,43,1891,40379,9444811,55866875)
- 27: (2,3,7,43,1943,25615,456729463,450222796871)
- 28: (2,3,7,43,1951,30571,118463,14484098803019)
- 29: (2,3,7,43,2105,12773,2775277,168100338289)
- 30: (2,3,7,43,2137,16921,37501,49708999789)
- 31: (2,3,7,43,2755,5407,172771,357538828973647)
- 32: (2,3,7,43,2813,5045,692705317,188433744928309)

- 33: (2,3,7,43,3263,4051,2558951,61088439723561979)
- 34: (2,3,7,43,3559,3667,33816127,797040720326433787)
- 35: (2,3,7,47,395,779731,607979652631,369639258012703445569531)
- 36: (2,3,7,47,395,779731,607979652647,21743485766025360000683)
- 37: (2,3,7,47,395,779731,607979652683,6974325623477705424647)
- 38: (2,3,7,47,395,779731,607979653531,410254449012081168631)
- 39: (2,3,7,47,395,779731,607979655287,139119028839856004123)
- 40: (2,3,7,47,395,779731,607979697799,8183472856913555659)
- 41: (2,3,7,47,395,779731,607979793451,2624887933109395111)
- 42: (2,3,7,47,395,779731,607982046587,154405744751990423)
- 43: (2,3,7,47,395,779743,46768385339,1672627310178141725483)
- 44: (2,3,7,47,395,779747,35764242947,12154487527525118239)
- 45: (2,3,7,47,395,779827,6286857907,2158880732959)
- 46: (2,3,7,47,395,779831,6020372531,3660733426607933569531)
- 47: (2,3,7,47,395,781727,305967719,125881309327)
- 48: (2,3,7,47,395,782111,257276179,57278664659)
- 49: (2,3,7,47,395,782287,277442411,1701723083)
- 50: (2,3,7,47,395,782611,211810259,1592773460578079)
- 51: (2,3,7,47,395,816247,17428931,652510750371360683)
- 52: (2,3,7,47,395,1108727,2627707,140495574531059)
- 53: (2,3,7,47,403,19403,15435513367,238255072887400163323)
- 54: (2,3,7,47,403,19403,15435513395,8215692183434294399)
- 55: (2,3,7,47,403,19403,15435513463,2456237880094942747)
- 56: (2,3,7,47,403,19403,15435516179,84697872837562655)
- 57: (2,3,7,47,415,81111,6644612311,44150872756848148411)
- 58: (2,3,7,47,415,81111,6644612339,1522443894582665279)
- 59: (2,3,7,47,415,81111,6644613463,38292177286592827)
- 60: (2,3,7,47,415,81111,6644645747,1320426321921983)
- 61: (2,3,7,47,449,4477,12137,34035763385)
- 62: (2,3,7,47,583,1223,1407479767,1980999293106894523)
- 63: (2,3,7,47,583,1223,1407479807,48317057302587443)
- 64: (2,3,7,47,583,1223,1468268915,33995520959)
- 65: (2,3,7,47,583,1223,2202310039,3899834875)
- 66: (2,3,7,53,209,10589,19651,86321)
- 67: (2,3,7,53,269,817,7301713,48932949591475)
- 68: (2,3,7,53,401,409,351691,397617853)
- 69: (2,3,7,55,179,24323,10057317271,101149630679497570171)
- 70: (2,3,7,55,179,24323,10057317287,5949978284730273323)
- 71: (2,3,7,55,179,24323,10057317311,2467064172726591731)
- 72: (2,3,7,55,179,24323,10057317467,513449911932648503)
- 73: (2,3,7,55,179,24323,10057317967,145121431390804003)
- 74: (2,3,7,55,179,24323,10057320619,30202945461748519)
- 75: (2,3,7,55,179,24323,10057325347,12523178395739983)
- 76: (2,3,7,55,179,24323,10057454579,736667018400959)
- 77: (2,3,7,61,187,485,150809,971259409)
- 78: (2,3,7,61,293,457,551,21709309)
- 79: (2,3,7,65,121,6271,1579937,2869621)
- 80: (2,3,7,67,113,28925,48220169,4074021053)
- 81: (2,3,7,67,113,29153,3712777,45401353)

- 82: (2,3,7,67,113,34477,178945,178344158228021)
- 83: (2,3,7,67,187,283,334651,49836124516795)
- 84: (2,3,7,71,103,65059,1101031,4400294969594807)
- 85: (2,3,11,17,79,301,1049,3696653)
- 86: (2,3,11,17,97,151,444161,317361415625)
- 87: (2,3,11,17,101,149,3109,52495396603)
- 88: (2,3,11,23,31,47059,2214502423,4904020979258368507)
- 89: (2,3,11,23,31,47059,2214502427,980804197623275639)
- 90: (2,3,11,23,31,47059,2214502475,92528699894575367)
- 91: (2,3,11,23,31,47059,2214502687,18505741750517011)
- 92: (2,3,11,23,31,47059,2214502831,11990273552017987)
- 93: (2,3,11,23,31,47059,2214504467,2398056482005535)
- 94: (2,3,11,23,31,47059,2214524099,226233749172527)
- 95: (2,3,11,23,31,47059,2214610807,45248521436443)
- 96: (2,3,11,23,31,47059,2215070383,8636647107907)
- 97: (2,3,11,23,31,47059,2217342227,1729101023519)
- 98: (2,3,11,23,31,47059,2244604355,165128325167)
- 99: (2,3,11,23,31,47059,2294166883,63772955407)
- 100: (2,3,11,23,31,47059,2365012087,34797266971)
- 101: (2,3,11,23,31,47059,2446798471,23325584587)
- 102: (2,3,11,23,31,47059,2612824727,14526193019)
- 103: (2,3,11,23,31,47059,3375982667,6436718855)
- 104: (2,3,11,23,31,47063,442938131,980970939025927675)
- 105: (2,3,11,23,31,47063,447473399,43702604167)
- 106: (2,3,11,23,31,47095,59897203,132743972247361531)
- 107: (2,3,11,23,31,47119,36349891,4619150372467)
- 108: (2,3,11,23,31,47131,30382063,67384091875543675)
- 109: (2,3,11,23,31,47147,24928579,11061526082145911)
- 110: (2,3,11,23,31,47243,12017087,26715920281613179)
- 111: (2,3,11,23,31,47423,6114059,13644326865136507)
- 112: (2,3,11,23,31,47479,5307047,2371471764522551)
- 113: (2,3,11,23,31,47491,5161279,4952592862147)
- 114: (2,3,11,23,31,49759,866923,2029951372029307)
- 115: (2,3,11,23,31,60563,211031,601432790177275)
- 116: (2,3,11,23,31,74963,126415,259118345891)
- 117: (2,3,11,23,31,84527,106159,84453127154999)
- 118: (2,3,11,25,29,787,264841,2542873)
- 119: (2,3,11,25,29,1097,2753,144508961851)
- 120: (2,3,11,31,35,67,369067,1770735487291)
- 121: (2,3,13,25,29,67,2981,11294561851)
- 122: (2,5,7,11,17,157,961,4398619)

**Appendix B: All solutions of  $\sum_{1 \leq i \leq N} \frac{1}{x_i} - \prod_{1 \leq i \leq N} \frac{1}{x_i} = 1$  for  $N \leq 7$**

- 1: (2,3,5)
- 2: (2,3,11,13)
- 3: (2,3,7,41)
- 4: (2,3,11,17,59)
- 5: (2,3,7,83,85)
- 6: (2,3,7,43,1805)
- 7: (2,3,11,23,31,47057)
- 8: (2,3,7,71,103,61429)
- 9: (2,3,7,53,271,799)
- 10: (2,3,7,47,481,2203)
- 11: (2,3,7,47,395,779729)
- 12: (2,3,7,43,3611,3613)
- 13: (2,3,7,43,3041,4447)
- 14: (2,3,7,43,2501,6499)
- 15: (2,3,7,43,2167,10841)
- 16: (2,3,7,43,2053,15011)
- 17: (2,3,7,43,1945,25271)
- 18: (2,3,7,43,1901,36139)
- 19: (2,3,7,43,1871,51985)
- 20: (2,3,7,43,1825,173471)
- 21: (2,3,7,43,1819,252701)
- 22: (2,3,7,43,1811,654133)
- 23: (2,3,7,43,1807,3263441)
- 24: (2,3,11,23,31,94115,94117)
- 25: (2,3,11,23,31,47059,2214502421)
- 26: (2,3,7,71,103,67213,713863)
- 27: (2,3,7,71,103,62857,2704339)
- 28: (2,3,7,71,103,61955,7238201)
- 29: (2,3,7,71,103,61559,29133437)
- 30: (2,3,7,71,103,61477,79005919)
- 31: (2,3,7,71,103,61441,319853515)
- 32: (2,3,7,59,163,1381,775807)
- 33: (2,3,7,55,179,24323,10057317269)
- 34: (2,3,7,47,583,1223,1407479765)
- 35: (2,3,7,47,449,3299,379591)
- 36: (2,3,7,47,415,8111,6644612309)
- 37: (2,3,7,47,403,19403,15435513365)
- 38: (2,3,7,47,401,25535,1837531099)
- 39: (2,3,7,47,395,779731,607979652629)
- 40: (2,3,7,47,395,788491,70175789)
- 41: (2,3,7,47,395,779819,6832003021)
- 42: (2,3,7,47,395,1559459,1559461)
- 43: (2,3,7,43,3307,3979,642279641)
- 44: (2,3,7,43,2533,7807,32435)

- 45: (2,3,7,43,2159,11047,98567401)
- 46: (2,3,7,43,1907,43115,163073)
- 47: (2,3,7,43,1907,34165,17766223)
- 48: (2,3,7,43,1823,193667,637617223445)
- 49: (2,3,7,43,1807,3263443,10650056950805)
- 50: (2,3,7,43,1807,6526883,6526885)



# Appendix C: All 550 solutions of $\sum_{1 \leq i \leq 8} \frac{1}{x_i} - \prod_{1 \leq i \leq 8} \frac{1}{x_i} = 1$

- 1: (2,3,7,43,1807,3263443,10650056950807,113423713055421844361000441)
- 2: (2,3,7,43,1807,3263443,10650056950811,22684742611092888917760733)
- 3: (2,3,7,43,1807,3263443,10650056950847,2766432025752386503391041)
- 4: (2,3,7,43,1807,3263443,10650056950895,1274423742206281453625521)
- 5: (2,3,7,43,1807,3263443,10650056951011,553286405158997346238853)
- 6: (2,3,7,43,1807,3263443,10650056951251,254884748449776336285749)
- 7: (2,3,7,43,1807,3263443,10650056951353,207355965377032496263511)
- 8: (2,3,7,43,1807,3263443,10650056951413,186859494335874429774611)
- 9: (2,3,7,43,1807,3263443,10650056951839,109800303065259151146401)
- 10: (2,3,7,43,1807,3263443,10650056953541,41471193083926544813347)
- 11: (2,3,7,43,1807,3263443,10650056953841,37371898875694931515567)
- 12: (2,3,7,43,1807,3263443,10650056954455,31083505917860578820921)
- 13: (2,3,7,43,1807,3263443,10650056955925,22157396583303152145971)
- 14: (2,3,7,43,1807,3263443,10650056955971,21960060621571875789925)
- 15: (2,3,7,43,1807,3263443,10650056969051,6216701192092161324829)
- 16: (2,3,7,43,1807,3263443,10650056973233,5057462580561823763311)
- 17: (2,3,7,43,1807,3263443,10650056975693,4557548652728700190411)
- 18: (2,3,7,43,1807,3263443,10650056976401,4431479325180675989839)
- 19: (2,3,7,43,1807,3263443,10650056981857,3652819986026411796191)
- 20: (2,3,7,43,1807,3263443,10650056993159,2678056182713693394601)
- 21: (2,3,7,43,1807,3263443,10650056999489,2329842318137500089151)
- 22: (2,3,7,43,1807,3263443,10650057004829,2099544890708757769051)
- 23: (2,3,7,43,1807,3263443,10650057042743,1233711280926563627161)
- 24: (2,3,7,43,1807,3263443,10650057062941,1011492524632410313307)
- 25: (2,3,7,43,1807,3263443,10650057064837,994674380386454897891)
- 26: (2,3,7,43,1807,3263443,10650057075241,911509739065785598727)
- 27: (2,3,7,43,1807,3263443,10650057106061,730564005725327919883)
- 28: (2,3,7,43,1807,3263443,10650057160685,540424317300132443371)
- 29: (2,3,7,43,1807,3263443,10650057162571,535611245062784239565)
- 30: (2,3,7,43,1807,3263443,10650057194221,465968472147545578475)
- 31: (2,3,7,43,1807,3263443,10650057220921,419908986661797114455)
- 32: (2,3,7,43,1807,3263443,10650057282835,341607861336024725621)
- 33: (2,3,7,43,1807,3263443,10650057406397,248959522702338919291)
- 34: (2,3,7,43,1807,3263443,10650057410491,246742264705358286077)
- 35: (2,3,7,43,1807,3263443,10650057515857,200731826106380706191)
- 36: (2,3,7,43,1807,3263443,10650057520961,198934884597336540223)
- 37: (2,3,7,43,1807,3263443,10650057577837,180890131003613942891)
- 38: (2,3,7,43,1807,3263443,10650058000201,108084871980072049319)
- 39: (2,3,7,43,1807,3263443,10650058223897,89093180781187556791)
- 40: (2,3,7,43,1807,3263443,10650058610951,68321580787250505769)
- 41: (2,3,7,43,1807,3263443,10650058946809,56825432783897027351)
- 42: (2,3,7,43,1807,3263443,10650059165749,51208422358805751251)
- 43: (2,3,7,43,1807,3263443,10650059228761,49791913060513344503)
- 44: (2,3,7,43,1807,3263443,10650059714345,41042931721701387271)
- 45: (2,3,7,43,1807,3263443,10650059750899,40507134183243596501)

- 46: (2,3,7,43,1807,3263443,10650059776061,40146373741321701883)
- 47: (2,3,7,43,1807,3263443,10650060058039,36503135151956613401)
- 48: (2,3,7,43,1807,3263443,10650060085961,36178034720768349223)
- 49: (2,3,7,43,1807,3263443,10650060720223,30090529437288820961)
- 50: (2,3,7,43,1807,3263443,10650061626077,24260361131432510491)
- 51: (2,3,7,43,1807,3263443,10650062238733,21449571707804380811)
- 52: (2,3,7,43,1807,3263443,10650063316261,17818644676283072003)
- 53: (2,3,7,43,1807,3263443,10650066930821,11365095076824966115)
- 54: (2,3,7,43,1807,3263443,10650067099565,11176127163949062571)
- 55: (2,3,7,43,1807,3263443,10650068025521,10241692991806710895)
- 56: (2,3,7,43,1807,3263443,10650070563995,8331909447275677021)
- 57: (2,3,7,43,1807,3263443,10650070768501,8208594864385838099)
- 58: (2,3,7,43,1807,3263443,10650070951271,8101435356694279945)
- 59: (2,3,7,43,1807,3263443,10650072486971,7300635550436883325)
- 60: (2,3,7,43,1807,3263443,10650073935703,6677926509977160761)
- 61: (2,3,7,43,1807,3263443,10650075630037,6072193870844315891)
- 62: (2,3,7,43,1807,3263443,10650075797891,6018114407503324837)
- 63: (2,3,7,43,1807,3263443,10650075798763,6017835980166266861)
- 64: (2,3,7,43,1807,3263443,10650080117897,4895908588016066791)
- 65: (2,3,7,43,1807,3263443,10650080327161,4852080746332062743)
- 66: (2,3,7,43,1807,3263443,10650082659077,4411964805021755491)
- 67: (2,3,7,43,1807,3263443,10650083390441,4289922861606436807)
- 68: (2,3,7,43,1807,3263443,10650086501387,3838301107202656141)
- 69: (2,3,7,43,1807,3263443,10650089026489,3536138409375784151)
- 70: (2,3,7,43,1807,3263443,10650107240345,2255424306869577271)
- 71: (2,3,7,43,1807,3263443,10650107694601,2235233952835373159)
- 72: (2,3,7,43,1807,3263443,10650112756565,2032483912456467571)
- 73: (2,3,7,43,1807,3263443,10650119325763,1818428145004661861)
- 74: (2,3,7,43,1807,3263443,10650125016751,1666390409500696049)
- 75: (2,3,7,43,1807,3263443,10650126167623,1638683417332729961)
- 76: (2,3,7,43,1807,3263443,10650141875291,1335593822040992797)
- 77: (2,3,7,43,1807,3263443,10650150346961,1214447294214423823)
- 78: (2,3,7,43,1807,3263443,10650151190591,1203575716078814017)
- 79: (2,3,7,43,1807,3263443,10650170255905,1001057505462912671)
- 80: (2,3,7,43,1807,3263443,10650171754619,987989272817600701)
- 81: (2,3,7,43,1807,3263443,10650172786261,979190237648774003)
- 82: (2,3,7,43,1807,3263443,10650174744829,962909362289669051)
- 83: (2,3,7,43,1807,3263443,10650184347359,890330759859381601)
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# New Directions of Modern Cryptography

Modern cryptography has evolved dramatically since the 1970s. With the rise of new network architectures and services, the field encompasses much more than traditional communication where each side is of a single user. It also covers emerging communication where at least one side is of multiple users. **New Directions of Modern Cryptography** presents general principles and application paradigms critical to the future of this field.

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